THESIS FOR THE DEGREE OF LICENTIATE OF PHILOSOPHY

Unified Static and Runtime Verification of Object-Oriented Software

JESÚS MAURICIO CHIMENTO

CHALMERS | GÖTEBORG UNIVERSITY

Department of Computer Science and Engineering
CHALMERS UNIVERSITY OF TECHNOLOGY AND GÖTEBORG UNIVERSITY

Göteborg, Sweden 2016

Unified Static and Runtime Verification of Object-Oriented Software
JESÚS MAURICIO CHIMENTO

Technical Report 162L
ISSN 1652-876X
Department of Computer Science and Engineering
Research group: Formal Methods

Department of Computer Science and Engineering
CHALMERS UNIVERSITY OF TECHNOLOGY and GÖTEBORG UNIVERSITY
SE-412 96 Göteborg
Sweden
Telephone +46 (0)31-772 1000

Printed at Chalmers
Göteborg, Sweden 2016

Abstract

At the time of verifying software one can make use of several verification techniques. These techniques mostly fall in one of two categories: *Static Verification* and *Dynamic Verification*. *Runtime Verification* is a dynamic verification technique which is concerned with the monitoring of software, providing guarantees that observed runs comply with specified properties. It is strong in analysing systems of a complexity that is difficult to address by static verification, e.g., systems with numerous interacting sub-units, real (as opposed to abstract) data, etc. On the other hand, the major drawbacks of runtime verification are the impossibility to extrapolate correct observations to all possible executions, and that the monitoring of a program introduces runtime overheads.

The work presented in this thesis addresses these issues by introducing a novel approach which combines the use of runtime verification with static verification, in such a way that: (i) static verification attempts to 'resolve' the parts of the properties which can be confirmed statically; (ii) the static results, even if only partial, are used to improve the specified properties such that generated monitors will not check at runtime what was already verified statically.

In addition, this thesis introduces the specification language *ppDATE* (and its semantics), which allows to describe properties suitable for static and runtime verification within a single formalism; the verification tool *StaRVOOrS*, which embodies the previously mentioned approach; and presents some case studies to demonstrate the effectiveness of using this new approach.

Acknowledgments

The cover of this thesis has only my name on it, but it would not have been possible to finish it without the support and collaboration of others.

First of all, I would like to express my sincere gratitude to both my supervisor Wolfgang Ahrendt, and my co-supervisor Gerardo Schneider, for their guidance, their valuable insights, and their advice during my first steps towards my PhD. It has been a pleasure working with them, and I am looking forward to continue doing so in the years to come.

Many thanks to Gordon Pace, Christian Colombo, Richard Bubel, and Martin Henschel for their collaboration. Completing this thesis without their comments and help would have been much more harder for me.

I would also like to thank my collegues, but most important friends, at Chalmers: Raúl, Danielito, John, Evgeny, Pablo, Daniel, Hamid, Bart, Marco, and many others who I forgot to mention here. It is amazing going every day to work and always finding a terrific working environment there.

Thanks to the Happy people Malene, Stig, Maryam, Yolanda, Stefanie, Kengo, and Nina, for keeping life interesting outside work.

Ni que hablar de mis amigos en Argentina. Naza, Cartu, Fede, Ema, Taihu, Santiago, Eze, Chacho, Fer, Tincho, Exe, Gustavo, gracias muchachos por estar siempre ahí cuando uno los necesita.

Last but not the least, gracias a mi familia por apoyarme siempre, y más ahora a la distancia. Siempre van a estar conmigo, no importa en donde esté.

CONTENTS

1 **Introduction** 1

2 **Combined Static and Runtime Verification of Data- and Control-Oriented Properties** 19

3 **StaRVOOrS User Manual** 103

CHAPTER
ONE

INTRODUCTION

Technology is getting more and more integrated into our ordinary activities. Desktop computers, laptops, netbooks, tablets, smart phones, and smart watches, represent just a short list of devices which are used on a daily basis all around the globe. Even though these devices may be really different, all of them have one thing in common: they all operate by executing programs (i.e., software).

Usually, software developers provide the users with an informal (general) description about what their programs are supposed to do. However, they may not offer any guarantees about the actual behaviour of their programs. In fact, it is quite a common practice for developers to include as part of the documentation of their programs a *Terms and Conditions* section, where they add disclaimers saying, for instance, that they do not take any responsibility if the use of their programs lead to a malfunction of the devices running them.

Unexpected program behaviours may be a real headache for their users. It is true that if a program which checks the weather forecast fails during its execution, it may not represent any harm. However, if the software in a self-driven car fails, it may be catastrophic for their users.

Fortunately, efforts by programmers to avoid unexpected behaviour on their programs is increasing. For instance, during the last decade developers (and the software industry) have started considering the use of *Formal Methods* to develop and verify their programs.

The main idea behind the use of formal methods is that, given the specification of the correctness properties of a program, one may use (formal) verification techniques in order to control whether the program fulfills its specification. Here, by specification of correctness properties we refer to a description of the behaviour of the program, e.g., under what conditions the program is intended to be executed, and which conditions are expected to hold once its execution is completed. Such specifications are usually written in a mathematical-based formalism.

Regarding the verification techniques, they may be divided into two categories: *static verification* techniques, and *dynamic verification* techniques.

On the one hand, *static verification* techniques deal with the analysis of either concrete source code, or a model of it. These techniques can verify properties over all possible executions of a program. However, as runtime data is not available during the static verification of properties, it might be necessary to introduce data abstractions to deal with a proof. This fact may make it difficult to achieved an automatic verification of the properties involving the use of such abstractions.

On the other hand, *dynamic verification* techniques are concerned with the monitoring of software, i.e., the program has to be executed in order to check the properties. These techniques, which in general are fully automated, provide guarantees that observed executions of a program comply to the specified properties. However, it is impossible to extrapolate correct observations to all possible executions. In addition, monitoring introduces runtime overheads which may be prohibitive in certain systems.

It is quite clear that static and dynamic verification have largely disjoint strengths. Therefore, combining both of them can allow the verification process to deal with richer properties, with a greater ease. This work presents a novel approach to address the verification of correctness properties, by combining the use of *runtime verification* (dynamic verification technique), and *deductive verification* (static verification technique).

The structure of this chapter is as follows: Section 1 introduces several verification techniques, including deductive verification and runtime verification. Section 2 introduces background information associated to the verification tools used in this thesis. Section 3 presents the project which was the context of this work, and elaborates on its objectives. Finally, Section 4 fully describes the contributions of this thesis, which are properly reflected on its different chapters.

1 Verification Techniques

The main objective behind the use of *verification techniques* is to check whether a program satisfies certain properties. Such properties usually describe the behaviour of the program under scrutiny. In general, these techniques are divided into either static or dynamic verification techniques, depending on whether the program under scrutiny has to be executed in order to verify it. Below, several of the most widely used static and dynamic verification techniques are briefly analysed. In addition, we make a (brief) general comparison between the use of both kind of techniques, and we discuss the benefits which may be obtained by combining them.

1.1 Static Verification Techniques

In this section we analyse two of the most widely used static verification techniques: *Deductive Verification*, and *Model Checking*.

Deductive Verification

Deductive verification [18] focuses on turning the correctness properties of a program into logical formulae, e.g., first-order logic, high-order logic, program logic, etc, and then verifying these formulae by deduction in a (logic) calculus.

In general, there are three main approaches that one may adopt to perform deductive verification. Let us call these three approaches *Proof Assistants*, *Program Logic*, and *Verification Condition Generation*.

Proof Assistants are interactive theorem provers which, in general, target some high-order logic. These provers are not language-oriented. Instead, they provide a language in which both the syntax and the semantics of the program under scrutiny have to be described. In addition, the correctness properties have to be modelled within the logic handled by the proof assistant. Then, one may use the proof assistant to develop the proof of the properties. Note that, even though proof assistants are interactive, they may present a certain degree of automation. As an example, the *Coq* [8] proof assistant targets *intuitionistic logic*, introduces the language *Gallina* to describe the syntax and the semantics of the program under scrutiny, and uses a *sequent calculus* to verify the correctness properties.

In relation to *Program Logic*, *Hoare Logic* [24] may be the most well-known program logic to analyse programs. Hoare logic offers both a clear notation to describe programs and their properties, and a set of axioms and inference rules which may be used to verify the properties [31]. In this logic, properties are described by using *Hoare triples*. A Hoare triple is simply an expression of the form $\{P\}\ S\ \{Q\}$, where S is the program under scrutiny, P is the precondition of the program, and Q is the postcondition of the program. In addition, one may consider the use of some language-oriented modal logic, e.g., *dynamic logic* [21], in order to reason about the correctness properties of the program under scrutiny. By using this kind of logic, the verification of the correctness properties may follow the flow of execution of the program under scrutiny. However, this may require the addition of some extra auxiliary specifications, e.g., loop invariants. As an example, one may refer to *KeY* [2]. This tool uses *Java dynamic logic* to deal with the correctness properties, and a *sequent calculus* to verify them.

On the *Verification Condition Generation* approach, programs are annotated with assertions representing the desired correctness properties. Then, this assertions may be used to generate first-order logic verification conditions which later may be discharged by using some automatic theorem prover. As an example, one may refer to *Dafny* [26]. Dafny is a programming language which natively supports specification constructs to describe the specification of the methods. Such specifications are used to generate first-order conditions which are discharged by using *Z3* [17].

Model Checking

Model Checking consists in verifying properties about a system by analysing a finite state abstraction of it, which is usually referred as *model* [11]. This technique determines whether a model fulfills the specified property by performing an exhaustive search over the entire state space of the system, aiming at finding an execution trace which violates it. If not such a trace exists, then the property is considered to be satisfied by the model.

As this technique deals with finite state systems, the tools implementing this technique, a.k.a. *model checkers*, automatically analyse all of the possible executions of the system, always terminating with a yes or no answer depending on whether the specification is fulfilled or not (if enough resources are available of course). However, when dealing with a real world problem the model checkers may have to deal with the *state explosion problem* [12]. In short, when the amount of state variables in the system increases, the size of the system state space grows exponentially.

Whenever the specification is not fulfilled, i.e., a property is violated, the model checkers usually return the execution trace which has violated the property, usually referred as *error trace*. Such trace serves as a counter-example for the property.

As an example of a model checker one can refer to SPIN [25]. This tool is a popular model checker which uses the language *PROMELA* for the description of the models, and *Temporal Logic* as the specification language for the properties.

1.2 Dynamic Verification Techniques

In this section we analyse, possibly, the main two dynamic verification techniques: *Testing*, and *Runtime Verification*. Note that is this work we use the term dynamic verification to refer to any of the previous techniques, and not as a synonym for runtime verification.

Testing

Testing [23] aims at analysing particular executions of a program to determine whether they produce certain expected values, i.e., the program behaves as expected. Such executions are based on *test cases*. A test case represents a possible initial state of the program under scrutiny, and the expected state of the program once its execution is completed. It is defined by assigning particular values to the different variables and parameters associated to the program.

In general, testing is divided into two categories: *Black-box testing* and *White-box testing*.

Black-box testing focuses on the analysis of the functionality of the program under scrutiny, treating it as a '*black box*', i.e., without looking into its source code. Therefore, it is not necessary for the person testing the program to know exactly how it is structured, but only to have some idea of what the program is supposed to do. For instance, if the program is sorting an array, then one only has to know that the array has to be sorted after executing the program, without the need to know which sorting algorithm the program is implementing. In addition, this kind of testing is quite useful whenever (part of) the source code of the program is not available. Among the different black-box testing techniques, *Input Space Partitioning* is usually one the most highlighted ones. This technique consists in, first, defining the *input space* of the program, i.e., the set of all possible inputs that may be fed to the program. Next, the input space is divided into several *disjoint partitions* such that the values on each partition test different functionalities of the program. Then, test cases are generated by selecting a value from each partition. Finally, these test cases are used to execute the program, and the obtained results are analysed.

Regarding *White-box testing*, it analyses the structure of a program to trace possible execution paths through its code. Therefore, one needs to have access to the complete source code of the program to perform this kind of testing. There are many criteria which can be followed when using this kind of testing such as, for instance, *Code Coverage* criteria. Within these criteria, one can highlight the *Statement Coverage* criterion, where all the test cases are generated in such a way that all the statements of a program are executed, or *Condition Coverage* criterion, where all the test cases are generated in such a way that every boolean expression of the program is evaluated to both true and false.

Runtime Verification

Runtime verification [22, 28] is concerned with the monitoring of software executions. This technique detects violations of properties which occur during the execution of the program under scrutiny. Due to this fact, runtime verification gives the possibility to react to incorrect behaviour of a program whenever an error is detected.

Properties to be verified using this technique are usually described in in two possible manners. One possibility is annotating the source code of the program under scrutiny with *assertions*. An assertion is a logical formula which is expected to be true whenever the execution of the annotated program reaches it. The other possibility is using a high level specification language. For instance, *Linear Temporal Logic* (LTL) [29] is one of the most popular formalisms in use. In addition, another approach which is increasing in popularity is writing properties using automaton-based specification languages [4, 14].

In order to verify properties, runtime verification introduces the use of *monitors*. A monitor is a piece of software which runs in parallel to the program under scrutiny, controlling that the execution of the program does not violate any property. In addition, monitors usually create a log file where they add entries reflecting the results which are obtained whenever they attempt to verify a property.

In general, monitors are automatically generated from either the annotated assertions, or the (high level) specification of the properties. In the case of the former, literature usually refers to this kind of runtime verification as *Runtime Assertion Checking*. As an example on how to apply runtime assertion checking, one may refer to *openJML* [13] or *jml4c* [32], which are tools that allow to perform runtime assertion checking over Java programs annotated with assertions written in the *Java Modelling Language* (JML) [27]. Regarding the latter, one can refer either to LARVA [15] or MarQ [30] as examples of tools which apply runtime verification by generating monitors from high level specification languages.

One downside on the use of runtime verification is that it introduces some overhead to the execution of the system. Thus, one of the main objectives of the developers of tools which use this technique is to reduce as much as possible such overhead.

1.3 Comparing Static and Dynamic Verification Techniques

It is quite clear that just by categorising the verification techniques into either static or dynamic, we are already comparing them with respect to whether or not the program is executed in a certain concrete state and environment. In addition, one may also consider their functionality to elaborate on a comparative analysis.

Static verification techniques are good to analyse the correctness of programs. These techniques can verify properties over all possible executions of a program or a model of it, and as they are used prior to the deployment of the programs, they do not affect the behaviour of the programs at runtime. These facts are advantages of the use of static techniques over the dynamic ones due to the fact the two of the major drawbacks in the use of dynamic verification techniques is that they cannot extrapolate the results of correct observations to all possible executions of a program, i.e., they verify only one execution at a time, and that the monitoring of properties introduces runtime overheads which can be prohibitive for certain systems, and may affect their execution.

On the other hand, the application of the static verification techniques usually is not quite straightforward. For instance, depending on the implementation of the methods, some annotations may have to be introduced in their source code in order to verify the properties. This would be the case of a method which has a loop in its implementation. Here, some loop invariants will have to be annotated in the source code of the method in order to verify the different properties related to it. In addition, the techniques working with abstractions from the real code may loose accuracy when they analyse a property involving the use of the (real) data of the system. These issues, may make it difficult for these techniques to achieve an automatic static verification of the properties. On the other hand, dynamic verification techniques work with the real data of the system when verifying its executions, and do not require the introduction of any extra notation in the source code of the programs.

Regarding the dynamic verification techniques, they are lightweight techniques which are usually strong in analysing programs of a complexity which is difficult to address by the static verification ones, like programs interacting with several other systems, heavy usage of mainstream (external) libraries, and real (as opposed to abstract) data. For instance, dynamic verification techniques can directly access the results of the calls to methods belonging to them, whereas the manner in which those results are computed is usually not accessible to the static ones. Thus, either some data abstractions may have to be introduced in many of the properties, or some specification has to be provided for the library, in order to verify the properties using static verification techniques.

Nonetheless, these techniques cannot be used to guarantee the correctness of a program, mainly due to the fact that, as opposed to the static verification techniques, they cannot extrapolate the results of correct observations to all possible executions.

1.4 On the Combination of Static and Dynamic Verification Techniques

Enhancing verification techniques by combining them with other verification techniques is a practice that is getting more and more attention. In this work we are mainly interested in the combination of static and dynamic verification techniques.

Combining static and dynamic verification techniques can allow the verification process to deal with properties with a greater ease. For instance, instead of possibly adding complicated abstractions to a property in order to statically handle the result of a call to a method belonging to an external library, one can attempt to verify such properties by using a dynamic verification technique that directly access the results of such method calls.

In addition, such combinations can introduce benefits regarding the verification performance. For instance, by using static verification techniques one could improve the performance of the dynamic ones by ignoring at runtime the verification of all the properties which were proved correct statically.

There are several possible technique combinations which could be analysed. The combination of testing and static verification techniques is one of the most explored ones, e.g. [5, 7, 16, 19, 20, 34]. Here, static verification can be used, for instance, to limit the dynamic efforts by filtering test cases, or to accomplish high coverage of the test cases. Another possibility would be the combination of runtime verification and static verification techniques. For instance, in [9] a static verification technique which reduces runtime instrumentation is used to improve the efficiency of runtime monitoring based on tracematches, and in [33] runtime verification is integrated with static code analysis in order to generate monitors which will allow to both check for possible faults in the system under scrutiny, and eliminate false positives obtained statically.

In particular, we will focus on the combination of runtime verification and deductive verification. Ideally, before using runtime verification, one may attempt to prove some of the properties in the specification of the program under scrutiny by using deductive verification. Then, all the properties proved correct statically can be removed from the specification. This would result in an improvement on the performance of the runtime monitoring, because the monitor generated using runtime verification would only focus on the properties which were not proved correct statically. In addition, those properties can be analysed with dynamic verification techniques in an attempt to identify why it was not possible to prove them, e.g., one can test whether a method is actually returning the result expected in its specification.

Note that the previous combination can be extended by analysing partial proofs, i.e., analysing the proofs of the properties which were not proved correct statically. For instance, the results obtained from the analysis of the partial proofs can be used to strengthen the properties in such a way that the dynamic verification techniques will verify at runtime only the parts of the properties which were not proved correct statically. This particular approach is addressed on this work. In Sec. 3 we elaborate on this.

2 Background

This section briefly introduces the different concepts which are used on this thesis. Sec. 2.1 introduces the selected tool, and its associated specification language, to perform deductive verification. Sec. 2.2 introduces

```
/*@ public normal_behaviour
  @ requires n > 0 && y > 0;
  @ ensures x == n && y > x;
  @*/
public void foo (int n) {
   x = n;
   y = y + x;
}
```

Fig. 1. JML specification for a particular Java method.

the selected tool, and its associated specification language, to perform runtime verification.

2.1 Deductive Verification using KeY

In this thesis we use the deductive verifier KeY [2], which given a Java program whose specification is annotated using *JML*, translates such annotations into *(Java) dynamic logic* formulae, and then attempts to verify them. Below, we briefly elaborate on the previously mentioned concepts.

JML

The *Java Modelling Language* (*JML*) [27] is a specification language which primarily focuses on the description of pre/post-conditions of methods and class invariants. This language is compatible with Java expression syntax, a fact that simplifies its use. Fig. 1 illustrates a JML specification (from line 1 to 5) for Java a method named `foo`. Line 1 describes which one is the behaviour expected for method `foo` (either normal as in this example, or exceptional); line 2 describes the precondition of `foo`; line 3 describes the postcondition of `foo`; and line 4 lists the variables of the class which are modified by executing `foo`.

Dynamic Logic

Dynamic logic [21] is a modal logic which is used to reason about programs. Due to the many differences between the different programming languages, it is not possible to have one single version of dynamic logic to analyse them all. Therefore, each language is usually associated to its own version of this logic. For instance, the version of dynamic logic used by KeY to analyse Java programs is referred as *Java dynamic logic* (or *Java DL*).

In particular, dynamic logic includes two modalities $[\ ]$ (referred as box), and $\langle\rangle$ (referred as diamond). Given a dynamic logic formula ϕ and a program p, $[p]\phi$ means that if p terminates its execution, then it is in a state where ϕ holds; and $\langle p\rangle\phi$ means that p terminates its execution and ϕ holds in the final state reached by p.

In addition, dynamic logic formulae are written using the traditional logical operators $\wedge, \vee, \rightarrow$, and $\neg$, and both universally and existentially quantifications over *logic variables*[1]. Note that as dynamic logic is specific for the programming language under scrutiny, its syntax to describe the programs will depend on the programming language in use. For instance, the following expressions are all examples of Java DL formulae:

(1) $[x := x * x\ ;]x > 0$
(2) $\langle x := x + y\ ;\rangle x > y$
(3) $\forall l \cdot l > 0 \rightarrow [x := l;]x > 0$
(4) $x > y \wedge y > 0 \rightarrow [y := y + x\ ;]y > x$

Note that the dynamic logic formula $\phi \rightarrow [p]\psi$ is valid if whenever formula ϕ holds, and the execution of p terminates, the formula ψ is fulfilled afterwards. Therefore, the previous formula could be regarded as the Hoare triple $\{\phi\}p\{\psi\}$.

KeY

KeY [2] is a deductive verification tool for data-centric functional correctness properties of Java programs which, given a Java program annotated with *JML*, generates proof obligations (i.e., formulae) in Java DL, and attempts to prove them. Fig. 2 roughly illustrates how KeY would generate the dynamic logic proof obligation in the column of the right, from the Java method `foo` which is annotated with JML in the column of the left.

Similarly to many other verification tools, KeY has a few restrictions: it does not support concurrency and floating-point arithmetic, and the generic types are expected to be compiled away. However, it is also worth mentioning that KeY fully covers Java Card, and that Java integer types, exceptions, and static initialization are accurately modelled on it [2].

At the core of KeY, a prover uses a *sequent calculus* to construct proof trees for the generated proof obligations, by following the *symbolic execution* paradigm. Let Γ be a set of formulae. Then, the expression $\Gamma \vdash \langle p \rangle \phi$ represents a sequent which holds if p starts in a state where all the formulae in Γ hold, and then it terminates in a state where ϕ holds. While p is verified, one could arrive at an intermediate proof node. Such a node will look like as follows: $\Gamma \vdash \sigma\langle p' \rangle \phi$. This sequent means that, if Γ was fulfilled before p, and σ is the accumulated effect up to now, then ϕ will hold after executing the remaining program p'. Note that, in general, proofs may branch over case distinctions which are mainly triggered by Boolean decisions in the program. Such branching occurs by applying rules like the following simplified[2] `if` rule:

[1] Logic variables never occur in programs.
[2] The simplified rule ignores side effects or exceptions possibly caused by b.

```
/*@ public normal_behaviour
  @ requires n > 0 && y > 0;
  @ ensures x == n && y > x;
  @*/
public void foo (int n) {
   x = n;
   y = y + x;
}
```

$$n > 0 \land y > 0 \rightarrow \\ \langle x := n \,; y := y + x; \rangle\; x = n \land y > x$$

Fig. 2. Rough example of a DL proof obligation generated by KeY.

$$\textsf{if}\; \frac{\Gamma, \sigma(b) \vdash \sigma\langle s_1 \quad \omega\rangle\phi \qquad \Gamma, \sigma(\neg b) \vdash \sigma\langle s_2 \quad \omega\rangle\phi}{\Gamma \vdash \sigma\langle \texttt{if}\ b\ s_1\ \texttt{else}\ s_2 \quad \omega\rangle\phi}$$

2.2 Runtime Verification using Larva

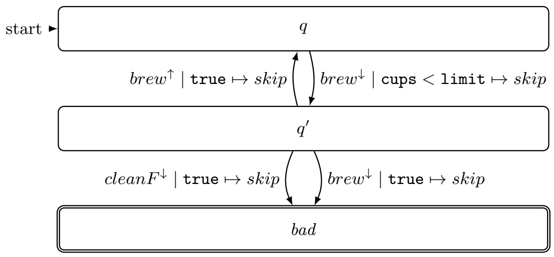
In this thesis we use the runtime verification tool Larva [15]. This tool automatically generates a runtime monitor from a property written in the automaton-based specification language *DATE* [14]. In order to do so, Larva transforms the set of properties into monitoring code together with AspectJ code, to link the system with the monitors.

Regarding *DATE*, it is a specification language to describe properties as finite state automata. Transitions in this language are tagged with labels of the form $e \mid cond \mapsto act$, where e represents a system event (primarily either an entry point $e^{\downarrow}$ or an exit point $e^{\uparrow}$ of a method), *cond* is a condition that must be true in order for the transition to take place, and *act* is is a code snippet to be performed when the transition is taken. Note that this description does not mention many of the features offered by *DATE*. One may refer to [14] for a more detail introduction to this language.

Fig. 3 illustrates an example of *DATE* describing a property about the system of a coffee machine. This *DATE* ensures that whenever the coffee machine is not active (i.e., is not brewing coffee) and the method `brew` starts the coffee brewing process, then it is not possible either to execute this method again or to execute the method `cleanF`, which initialises the task of cleaning the filter, until the current brewing process terminates.

3 StaRVOOrS: Towards the Combination of Runtime Verification with Static Verification

This thesis has been developed in the context of the research project *Unified Static and Runtime Verification of Object-Oriented Software*, or StaRVOOrS for short. This project, which is funded by *The Swedish*



start
q
$brew^{\uparrow} \mid \texttt{true} \mapsto skip$
$brew^{\downarrow} \mid \texttt{cups} < \texttt{limit} \mapsto skip$
q'
$cleanF^{\downarrow} \mid \texttt{true} \mapsto skip$
$brew^{\downarrow} \mid \texttt{true} \mapsto skip$
bad

Fig. 3. A *DATE* controlling the brew of coffee

Research Council (Vetenskapsrdet), has as a main purpose the development of a methodology for specifying and verifying both data- and control-oriented properties of object-oriented software systems, in a unified manner.

As a starting point to address this objective, Ahrendt et al. proposed in [6] a verification framework which combines the use of runtime verification with deductive verification. In short,

(i) Deductive verification is used to verify those parts of the properties which may be confirmed statically;
(ii) the previous results, even if they are only partial, are used to refine the original specification of the properties such that the monitors generated by using runtime verification will not have to check at runtime the statically verified parts of the properties. This reduces the overhead which the generated monitors add to the system.

In addition, [6] presents initial ideas about an automaton-based specification language, called *ppDATE*, which captures both the description of control-oriented and data-oriented properties. Basically, this language consists of a transition system alike the *DATE* formalism, whose states may include Hoare triples describing properties about the methods of the program under scrutiny.

One of the main contribution of this thesis consists in the full development of the syntax (both abstract and concrete), the grammar, and the formal semantics of *ppDATE*. Moreover, as a first approach on the use of the framework described above, this thesis introduces the combination of the deductive verifier KeY [2], and the runtime verifier Larva [15], in order to verify *Java* programs. Such combination is accomplished with the implementation of a verification tool which automatically combines the use both of the previous tools. Section 4 elaborates on these contributions.

4 Contributions of the Thesis

The contribuitions of this thesis have been disseminated in the following documents of two peer reviewed conference papers [4, 10], one peer reviewed journal paper (under submission) [3], and one user manual [1]. This section presents a brief description of each one of these works, and outlines the contributions of Mauricio Chimento in all of them.

Note that [3] and [1] correspond to the chapters 1 and 2 of this thesis, respectively. Regarding [4] and [10], as the content of [3] consists of an extension of both of them, we decided not to include them as part of this thesis. Anyhow, below we describe the contribuitions of Mauricio Chimento in these works. In addition, note that the format of all [1, 3] were adapted to suit the format of the thesis. However, their content remains unchanged.

4.1 A Specification Language for Static and Runtime Verification of Data and Control Properties

Paper [4] presents the development and formalisation of a notation language, called *ppDATE*, as an extension of the control-flow property language *DATE*, which is used in the runtime verification tool LARVA, and shows how specifications written in this notation can be analysed both using the deductive theorem prover KeY and the runtime verification tool LARVA. In addition, by using *ppDATE* to describe the corresponding specification, the STARVOORS verification framework is applied to Mondex, an electronic purse application.

Statement of contribution: The contributions of Mauricio Chimento on this paper are (i) collaborating on the formalisation of the *ppDATE* notation to describe the Hoare triples associated to the states of a *ppDATE*; (ii) formalising the Hoare triples which are part of the *ppDATE* specification for the Mondex case study; (iii) applying the STARVOORS verification framework to the Mondex case study; and (iv) performing some experiments to analyse (and compare) the overhead added to Mondex by the monitor generated using STARVOORS, and by the monitor that would be generated without using static verification to analyse the Hoare triples.

4.2 StaRVOOrS - A Tool for Combined Static and Runtime Verification of Java

Paper [10] presents the tool STARVOORS, which aims at both the specification and verification of properties by combining the use of runtime verification and static verification. This tool is fed with a Java program and a ppDATE specification of the program, and automatically generates a monitor in order to runtime verify the provided program. In order to do so, StaRVOOrS combines the deductive theorem prover KeY and the runtime verifier Larva. In addition, the effectiveness of this tool is demonstrated by applying it to Mondex, an electronic purse application.

Statement of contribution: The contributions of Mauricio Chimento on this paper are (i) a refinement of the original verification framework proposed in [6]; (ii) full development and implementation of the tool StaRVOOrS; (iii) and demonstrating the effectiveness of the tool by applying it to the Mondex case study.

4.3 Verifying Data- and Control-Oriented Properties Combining Static and Runtime Verification: Theory and Tools

The journal paper [3], which corresponds to the first chapter of this thesis, is basically an extension of the material presented in both [4] and [10]. In this paper, *ppDATE* is introduced as a proper specification language, and not just a simple notation. This is accomplished by introducing its syntax, grammar, and formal semantics. In addition, in order to cover new features of the *ppDATE* specification language, this paper introduces minor modifications into the algorithm used to translate a *ppDATE* specification into a *DATE* one, and shows the proof of correctness of such algorithm. Moreover, it shows the advantages of using the StaRVOOrS tool to verify a case study based on Mondex (electronic purse application), and a real world case study based on SoftSlate, an open-source Java shopping cart web application.

Statement of contribution: The contributions of Mauricio Chimento on this paper are (i) defining the syntax, the grammar (except the grammar of the templates), and the semantics (except the semantics of actions) of *ppDATE*; (ii) upgrading the translation algorithm to cover new features of the *ppDATE* specification language; (iii) proving the correctness of the translation algorithm; (iv) and using the StaRVOOrS tool to verify the SoftSlate case study. Note that in (i), Mauricio Chimento proposed initial versions for the definitions, and refined then with input from the co-authors. In addition, regarding (iii), Mauricio Chimento developed the whole proof of correctness on his own. However, some high level ideas related to the formalisation of the correctness theorem and its proof, in particular the introduction of the coupling invariants, were proposed by the co-authors.

4.4 StaRVOOrS User Manual

The StaRVOOrS user manual [1], which corresponds to the second chapter of this thesis, gives a high level explanation about how the StaRVOOrS tool works, provides an intuitive description of the *ppDATE* specification language, shows how to write a *ppDATE* specification in the input language of this tool by introducing its concrete grammar, and provides a complete example on how to use the tool.

Statement of contribution: The main contribution of Mauricio Chimento on this work is the introduction of the concrete grammar to write *ppDATE* specifications as a script, fact which is essential to use the tool StaRVOOrS.

References

1. StaRVOOrS User Manual. `cse-212294.cse.chalmers.se/starvoors/files/userguide.pdf`.
2. Wolfgang Ahrendt, Bernhard Beckert, Richard Bubel, Reiner Hähnle, Peter H. Schmitt, and Mattias Ulbrich, editors. *Deductive Software Verification—The KeY Book*, volume 10001 of *LNCS*. Springer, 2016.
3. Wolfgang Ahrendt, Jesús Mauricio Chimento, Gordon J. Pace, and Gerardo Schneider. Verifying data- and control-oriented properties combining static and runtime verification: Theory and tools (under submission). *Formal Methods in System Design*.
4. Wolfgang Ahrendt, Jesús Mauricio Chimento, Gordon J. Pace, and Gerardo Schneider. A specification language for static and runtime verification of data and control properties. In *FM'15*, volume 9109 of *LNCS*, pages 108–125. Springer, 2015.
5. Wolfgang Ahrendt, Christoph Gladisch, and Mihai Herda. Proof-based test case generation. In *Deductive Software Verification—The KeY Book*, volume 10001 of *LNCS*. Springer, 2016.
6. Wolfgang Ahrendt, Gordon Pace, and Gerardo Schneider. A Unified Approach for Static and Runtime Verification: Framework and Applications. In *ISOLA'12*, LNCS 7609. Springer, 2012.
7. Cyrille Artho and Armin Biere. Combined static and dynamic analysis. In *AIOOL'05*, volume 131 of *ENTCS*, pages 3–14, 2005.
8. Yves Bertot, Pierre Castran, Grard (informaticien) Huet, and Christine Paulin-Mohring. *Interactive theorem proving and program development : Coq'Art : the calculus of inductive constructions.* Texts in theoretical computer science. Springer, Berlin, New York, 2004.
9. Eric Bodden, Laurie J. Hendren, and Ondrej Lhoták. A staged static program analysis to improve the performance of runtime monitoring. In *ECOOP'07*, LNCS 4609, 2007.
10. Jesús Mauricio Chimento, Wolfgang Ahrendt, Gordon J. Pace, and Gerardo Schneider. StaRVOOrS: A Tool for Combined Static and Runtime Verification of Java. In Ezio Bartocci and Rupak Majumdar, editors, *Runtime Verification*, volume 9333 of *Lecture Notes in Computer Science*, pages 297–305. Springer International Publishing, 2015.

11. Edmund M. Clarke, Orna Grumberg, and Doron A. Peled. *Model checking.* MIT, Cambridge, Mass;London;, 1999.
12. Edmund M. Clarke, William Klieber, Miloš Nováček, and Paolo Zuliani. *Model Checking and the State Explosion Problem*, pages 1–30. Springer Berlin Heidelberg, Berlin, Heidelberg, 2012.
13. David R. Cok. *OpenJML: JML for Java 7 by Extending OpenJDK*, pages 472–479. Springer Berlin Heidelberg, 2011.
14. Christian Colombo, Gordon J. Pace, and Gerardo Schneider. Dynamic Event-Based Runtime Monitoring of Real-Time and Contextual Properties. In *FMICS'08*, volume 5596 of *LNCS*, pages 135–149. Springer-Verlag, September 2009.
15. Christian Colombo, Gordon J. Pace, and Gerardo Schneider. LARVA - A Tool for Runtime Monitoring of Java Programs. In *SEFM'09*, pages 33–37. IEEE Computer Society, 2009.
16. Christoph Csallner and Yannis Smaragdakis. Check 'n' crash: combining static checking and testing. In *27th International Conference on Software Engineering (ICSE 2005), 15-21 May 2005, St. Louis, Missouri, USA*, pages 422–431, 2005.
17. Leonardo M. de Moura and Nikolaj Bjrner. Z3: An efficient smt solver. In C. R. Ramakrishnan and Jakob Rehof, editors, *TACAS*, volume 4963 of *Lecture Notes in Computer Science*, pages 337–340. Springer, 2008.
18. Jean-Christophe Filliâtre. Deductive software verification. *International Journal on Software Tools for Technology Transfer*, 13(5):397–403, 2011.
19. Cormac Flanagan, K. Rustan M Leino, Mark Lillibridge, Greg Nelson, James B. Saxe, and Raymie Stata. Extended Static Checking for Java. In Jens Knoop and Laurie J. Hendren, editors, *PLDI'02*, pages 234–245. ACM, 2002.
20. Xi Ge, Kunal Taneja, Tao Xie, and Nikolai Tillmann. Dyta: dynamic symbolic execution guided with static verification results. In *Proceedings of the 33rd International Conference on Software Engineering, ICSE 2011, Waikiki, Honolulu , HI, USA, May 21-28, 2011*, pages 992–994, 2011.
21. David Harel, Dexter C. Kozen, and Jerzy Tiuryn. *Dynamic logic.* Foundations of computing. the MIT Press, Cambridge (Mass.), London, 2000.
22. Klaus Havelund and Grigore Roşu. Runtime verification. In *Computer Aided Verification (CAV'01) satellite workshop*, volume 55 of *ENTCS*, 2001.
23. Bill Hetzel. *The Complete Guide to Software Testing.* Wiley [Imprint], Hoboken, 2nd edition, 1993.
24. C. A. R. Hoare. An axiomatic basis for computer programming. *Commun. ACM*, 12(10):576–580, October 1969.
25. Gerard J. Holzmann. The model checker SPIN. *Software Engineering*, 23(5):279–295, 1997.
26. Jason Koenig and K. Rustan M. Leino. Getting Started with Dafny: A Guide. In Tobias Nipkow, Orna Grumberg, and Benedikt Hauptmann, editors, *Software Safety and Security*, volume 33 of *NATO Science for Peace and Security Series - D: Information and Communication Security*, pages 152–181. IOS Press, 2012.
27. Gary T. Leavens, Erik Poll, Curtis Clifton, Yoonsik Cheon, Clyde Ruby, David Cok, Peter Müller, Joseph Kiniry, and Patrice Chalin. *JML Reference Manual. Draft 1.200*, 2007.

28. Martin Leucker and Christian Schallhart. A Brief Account of Runtime Verification. *J. Log. Algebr. Program.*, 78(5):293–303, 2009.
29. Amir Pnueli. The temporal logic of programs. In *Proc. 18th IEEE Symposium on Foundation of Computer Science*, pages 46–57, 1977.
30. Giles Reger, Helena Cuenca Cruz, and David E. Rydeheard. MarQ: Monitoring at Runtime with QEA. In Christel Baier and Cesare Tinelli, editors, *TACAS*, volume 9035 of *Lecture Notes in Computer Science*, pages 596–610. Springer, 2015.
31. John C. Reynolds. *Theories of Programming Languages.* Cambridge University Press, 1st edition, 2009.
32. Amritam Sarcar and Yoonsik Cheon. A new eclipse-based jml compiler built using ast merging. Technical Report Technical Report UTEP-CS-10-08, University of Texas, 2010.
33. Hasan Sözer. Integrated static code analysis and runtime verification. *Softw., Pract. Exper.*, 45(10):1359–1373, 2015.
34. Julian Tschannen, Carlo A. Furia, Martin Nordio, and Bertrand Meyer. Usable verification of object-oriented programs by combining static and dynamic techniques. In Gilles Barthe, Alberto Pardo, and Gerardo Schneider, editors, *SEFM*, volume 7041 of *LNCS*, pages 382–398. Springer, 2011.

CHAPTER

TWO

COMBINED STATIC AND RUNTIME VERIFICATION OF DATA- AND CONTROL-ORIENTED PROPERTIES

Wolfgang Ahrendt, Jesús Mauricio Chimento, Gordon Pace, Gerardo Schneider

Abstract. Static verification is used to analyse and prove properties about programs before they are executed. Many of these techniques work directly on the source code, and are used to verify data-oriented properties over all possible executions. The analysis is necessarily an over-approximation as the real executions of the program are not available at analysis time. In contrast, runtime verification is usually more suitable for control-oriented properties, analysing the current execution path of the program in a fully automatic manner. In this article, we present a novel approach in which data-oriented and control-oriented properties may be stated in a single formalism amenable to both static and runtime verification. The specification language we present to achieve this is ppDATE, which enhances the language of DATE, with pre/postconditions. For runtime verification of ppDATE specifications, the language is translated into a DATE. We give a formal semantics to ppDATEs, using which we prove the correctness of our translation from ppDATEs to DATEs. We show how ppDATE specifications can be analysed using a combination of the deductive theorem prover KeY and the runtime verifier LARVA. Verification is performed in two steps: KeY first partially proves the data-oriented part of the specification, simplifying the specification which is then passed on to LARVA to check at runtime for the remaining parts of the specification including the control-oriented aspects. We show the applicability of our approach on two case studies.

1 Introduction

Runtime verification has been touted as a practical verification technique, and although it does not provide program analysis before deployment, it can check correct behaviour post-deployment by observing whether actual execution paths at runtime conform to the specification. Runtime verification scales up much more effectively than static analysis both in terms of performance and in terms of applicability to diverse contexts in which a program may interact with various other systems, services, libraries and be deployed.

Despite the fact that overheads induced by runtime verification might be small when compared to the computational effort required for static analysis, the fact that is done while the software is live can be problematic and prohibitive for certain systems. In this paper we present an approach to address the issue of runtime overheads through the use of static, deductive verification — an approach which also has the benefit of being able to verify parts of the specification a priori for all potential execution paths, leaving only parts which could not be proved before deployment to be checked dynamically.

Apart from the computational power required to perform the analysis, deductive and runtime verification have largely been applied to disjoint areas in which they are most effective — whereas deductive analysis excels in properties focusing on a system's data, runtime verification handles control-flow properties with substantially lower overheads than data-oriented ones. Combining the two approaches has the additional benefit that static analysis would typically be more effective in proving the parts of a specification which dynamic analysis would struggle most with. The challenge is thus to design a specification language which allows the expression of combined data- and control-flow properties in such a manner that they can be effectively decomposed for the application of different verification techniques.

The StaRVOOrS framework [8] addresses these issues by identifying a specification notation for such properties and a verification methodology combining static and dynamic analysis to verify combined control- and data-oriented properties. Although one may envisage different ways to combine static and dynamic analysis tools, a crucial requirement is that the specification languages used in the tools chosen are either identical, or can be somehow combined to allow for rich specifications getting the best of both approaches. Similar to *mode automata* [29] we have chosen to adopt an automata-based specification language (for the control-flow properties) but extended with data-flow properties encoded in the different states of the formalism.

This article is a significantly extended and revised version of two papers. In [6] we introduced the formalism *ppDATE*, where parts of the syntax where left underspecified, and we gave a high-level description of the algorithm to translate *ppDATE* into *DATE* [19], the formalism used in the runtime verification tool Larva [20]. In [17] we presented the tool StaRVOOrS, a full implementation of the framework introduced in [8, 6].

The novel contributions of this paper, going beyond the results reported in [6] and [17] are the following: i) We present a complete formal definition of *ppDATE* automata, including a formal semantics for the formalism (Sec. 5); ii) A proof of soundness of the algorithm to translate from *ppDATE* specifications into *DATE* ones (Sec. 7). iii) The application of our approach to SoftSlate Commerce, an open-source Java shopping cart web application (Sec. 9); iv) A description of the results of the case study including an analysis of the verification process providing evidence that our approach reduces the overhead of the runtime monitoring (Sec. 9).

2 Preliminaries

The work presented in this article is centred around static and runtime verification of Java systems. To implement these verification techniques, we use the deductive verifier KeY and the runtime verifier Larva. In this section, we introduce these tools at a high level of abstraction, but with sufficient detail to enable the understanding of the rest of the paper.

2.1 The deductive verifier KeY

KeY [5] is a deductive verification tool for data-centric *functional correctness* properties of Java source code. KeY generates proof obligations in *dynamic logic* (DL), a modal logic for reasoning about programs. DL extends first-order logic with two modalities, $\langle p \rangle \phi$ and $[p]\phi$, where p is a program and ϕ is another DL formula. The formula $\langle p \rangle \phi$ is true in a state s if there *exists* a terminating run of p, starting in s, resulting in a state where ϕ holds. The formula $[p]\phi$ holds in a state s if *all* terminating runs of p, starting in s, result in a state in which ϕ holds. For deterministic programs p, the only difference between the two modalities is that termination is *stated* in $\langle p \rangle \phi$, and *assumed* in $[p]\phi$.

KeY features (static) verification of Java source code annotated with specifications written in the *Java Modelling Language* (JML) [28]. JML allows for the specification of pre- and postconditions of method calls, and class/interface invariants. The main features of KeY are the translation of JML annotated Java programs to Java DL, and a theorem prover for validity of Java DL formulae, using a *sequent calculus*, covering almost all features of sequential Java (with the exception of generics and floating-point types currently). Given a set of formulae Γ, the sequent $\Gamma \vdash \langle p \rangle \phi$ holds if p, when starting in a state fulfilling all formulae in Γ, terminates in a state fulfilling ϕ. The calculus uses the *symbolic execution* paradigm.

For that, DL is extended by *explicit substitutions*. During the symbolic execution of p, the effects of p are gradually, starting from the front, turned into explicit substitutions. Thereby, after some proof steps, a certain prefix of p has turned into a substitution σ, representing the effects so far, while a remaining program p' is yet to be executed. While verifying p, an intermediate proof node may look like $\Gamma \vdash \sigma\langle p'\rangle\phi$. It tells us that, if Γ was true before the original program p, and σ is the accumulated effect up to now, then ϕ will be true after executing the remaining program p'.

As an example, consider a proof of the following DL sequent:

$$\texttt{x} > \texttt{0},\ \texttt{y} > \texttt{0}\ \vdash \langle\texttt{x=x+y;y=x-y;x=x-y;if(x\%2==0)\{}p_1\texttt{\}else\{}p_2\texttt{\};}q\rangle\phi \tag{1}$$

(where p_1, p_2, and q are Java fragments and ϕ is some postcondition). The sequent says that in each state where `x` and `y` are positive, the program given in the modality (which first swaps `x` and `y` using arithmetics) will terminate and result in a state where ϕ holds. When proving this sequent, the KeY prover will first, in a number of steps, turn the three leading assignments into explicit substitutions, apply the first to the second, the result to the third, and perform arithmetic simplification, arriving at

$$\texttt{x} > \texttt{0}, \texttt{y} > \texttt{0}\ \vdash (\texttt{x} \leftarrow \texttt{x+y} \,\|\, \texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})\langle\texttt{if(x\%2==0)\{}p_1\texttt{\}else\{}p_2\texttt{\};}q\rangle\phi$$

where $(\texttt{x} \leftarrow \texttt{x+y} \,\|\, \texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})$ denotes the explicit (parallel) substitution resulting from symbolic execution of the first three statements. A 'right-win' semantics is adopted to resolve clashes in substitutions, such that the above simplifies to:

$$\texttt{x} > \texttt{0},\ \texttt{y} > \texttt{0}\ \vdash (\texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})\langle\texttt{if(x\%2==0)\{}p_1\texttt{\}else\{}p_2\texttt{\};}q\rangle\phi$$

In general, most proofs branch over case distinctions, often triggered by Boolean decisions in the source code. The branching happens by applying rules like the following, simplified[1] if rule:

$$\mathsf{if}\ \frac{\Gamma, \sigma(b) \vdash \sigma\langle s_1\ \ \omega\rangle\phi \qquad \Gamma, \sigma(\neg b) \vdash \sigma\langle s_2\ \ \omega\rangle\phi}{\Gamma \vdash \sigma\langle\texttt{if}\ b\ s_1\ \texttt{else}\ s_2\ \ \omega\rangle\phi}$$

In our example, applying the if rule to the latest sequent results in splitting the proof into two branches, with the following sequents, respectively:

$$\texttt{x} > \texttt{0},\ \texttt{y} > \texttt{0},\ (\texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})(\texttt{x\%2} = \texttt{0})\ \vdash (\texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})\langle p_1\texttt{;}q\rangle\phi$$
$$\texttt{x} > \texttt{0},\ \texttt{y} > \texttt{0},\ (\texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})(\neg(\texttt{x\%2} = \texttt{0}))\ \vdash (\texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})\langle p_2\texttt{;}q\rangle\phi$$

Applying the substitution on the left side of either sequent results in:

$$\texttt{x} > \texttt{0},\ \texttt{y} > \texttt{0},\ \texttt{y\%2} = \texttt{0}\ \vdash (\texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})\langle p_1\texttt{;}q\rangle\phi \tag{2}$$
$$\texttt{x} > \texttt{0},\ \texttt{y} > \texttt{0},\ \neg(\texttt{y\%2} = \texttt{0})\ \vdash (\texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})\langle p_2\texttt{;}q\rangle\phi \tag{3}$$

[1] The simplified rule ignores side effects or exceptions possibly caused by b.

Note that in this step, by applying the swapping substitution, the branching condition (`x` being even or odd) on the *state after swapping* got translated into a condition on the *prestate* of the original program p, before the swapping. The resulting sequents tell us, among other things, that if `y` is even (respectively odd) in the prestate of p, then path p_1 (respectively p_2) is taken in the execution of p. In general, when building a proof in such a symbolic manner, the left side of sequents accumulate conditions on the original prestate through a particular execution path.

Once all proof branches are closed, we have a *complete proof* of the root sequent. However, a proof attempt may result in a *partial proof*, only, where some proof branches are closed and others are not. Such partial proofs are important for the work presented in this article. In the above example, consider a partial proof where the left branch, i.e., the sub-proof for sequent (2), is closed, whereas the right branch, i.e., the sub-proof for sequent (3), is *not* closed. From this partial proof, we can conclude that the following modification of the root sequent (1) is valid:

$$\texttt{x} > \texttt{0}, \texttt{y} > \texttt{0}, \texttt{y\%2} = \texttt{0} \;\vdash\; \langle \texttt{x=x+y;y=x-y;x=x-y;if(x\%2==0)\{}p_1\texttt{\}else\{}p_2\texttt{\};}q\rangle\phi \tag{4}$$

(We added $\texttt{y\%2} = \texttt{0}$ to the left side of (1), as additional assumption.) This sequent can be proven by replaying the original proof, where now both branches would close. The left branch closes as the sub-proof for (2) will replay identically. The right branch closes because the following variant of (3) can be closed immediately, due to contradicting assumptions:

$$\texttt{x} > \texttt{0},\ \texttt{y} > \texttt{0},\ \texttt{y\%2} = \texttt{0},\ \neg(\texttt{y\%2} = \texttt{0}) \;\vdash\; (\texttt{y} \leftarrow \texttt{x} \,\|\, \texttt{x} \leftarrow \texttt{y})\langle p_2\texttt{;}q\rangle\phi$$

2.2 The runtime verifier Larva

Larva[2] [20] is an automata-based runtime verification tool for Java programs. As with many other runtime verifiers, Larva automatically generates a runtime monitor from a property written in a formal language, in its case using *Dynamic Automata with Timers and Events* (DATEs). Transitions in a *DATE* are of the form: *event* | *condition* $\mapsto$ *action*, where *event* is what triggers the transition, the *condition* is checked and must hold in order the transition to take place, and the *action* is a code snippet to be performed when taking the transition (after checking the condition). DATEs are an extension of timed automata — they are effectively finite state automata, whose transitions are triggered by system events (primarily entry points $\texttt{f}^{\downarrow}$ and exit points $\texttt{f}^{\uparrow}$ of methods) and timers, but augmented with: (i) A symbolic state which may be used as conditions to guard transitions and can be modified via actions also specified on the transition; (ii) replication of automata, through which a new automaton is created for each discovered instance of an object; (iii) communication between

[2] *Logical Automata for Runtime Verification and Analysis.*

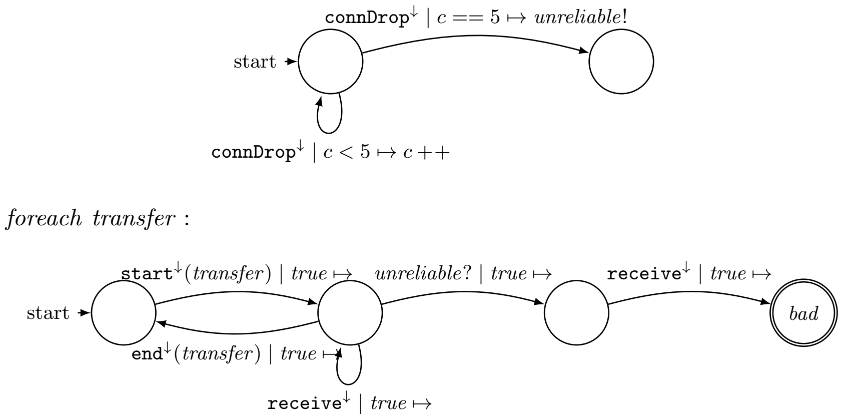


Fig. 1. Example of a *DATE* specification.

automata using standard CCS-like channels with $c!$ acting as a broadcast on channel c and which can be read by another automaton matching on event $c?$. Full details of the formalisation of DATEs can be found in [20].

The automata illustrated in Fig. 1 represent an example of *DATE* automata describing a property which should hold during a connection. The first automaton ensures that if the connection drops (event $\texttt{connDrop}^{\downarrow}$) occurs five times, a message is broadcast (over channel *unreliable*) to highlight the fact that the connection port is unreliable. The second automaton (with the *foreach* keyword) ensures that every time a file transfer is initiated, an automaton is created to monitor that transfer. If during the transfer (i.e. between the events $\texttt{start}^{\downarrow}$ and $\texttt{end}^{\downarrow}$) one receives event *unreliable*?, no further transfers may occur.

In order to monitor a system using Larva, the user must provide the system to be monitored (a Java program) and a set of properties in the form of a Larva script (a textual representation of DATEs). Larva transforms the set of properties into monitoring code together with AspectJ code to link the system with the monitors. Since the Java byte code is used for instrumentation, it is possible to monitor third-party software with Larva, though knowledge of methods names is still required.

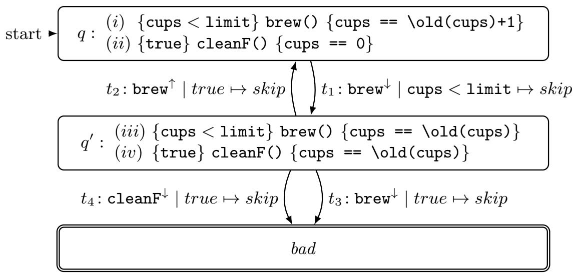
3 *ppDATE*: A Specification Language for Data- and Control-oriented Properties

In many cases, verification tools perform more effectively on a particular styles of specification. In combining two different verification tools which use very different analysis techniques, one challenge is that if we adopt an off-the-shelf language, we cannot expect to derive useful verification results from both tools. Given that deductive verification tools like KeY perform much better on data-centric properties, while runtime verification tools like Larva perform better on control-flow properties, we have defined a specification language to combine the two types of properties. In real scenarios, there is often a need to specify both, rich data constraints and legal execution sequences.

Data-oriented properties are typically written in expressive formalisms (like first-order logic), but typically give invariants about specific points in the execution of a system, rather than properties across traces of execution. The *Java Modelling Language* (JML) [28] is one such language, which focuses primarily on pre/postconditions of method calls and class invariants, but is not well suited for specifying which sequences of events or states are correct. In contrast, *control-oriented* specification languages specialise primarily on identifying legal sequences of events or states, for instance using automata or temporal logics. Although constraints about the data are possible, they are usually cumbersome and greatly increase the computational complexity required to verify them. *Dynamic Automata with Timers and Events* (*DATE*) [19] are one such specification language.

Coding control-flow into data-centric languages, like coding legal execution traces via model/ghost fields in JML, or including data-flow information in control-centric languages, like considering variable updates as events in *DATE* specification, can lead to substantial increase in the complexity of the specification from an understandability and/or verification perspective.

In order to address this, we propose *ppDATE*, a formalism to deal with both types of properties ensuring understandability and tractability of analysis using the StaRVOOrS verification framework. *ppDATE* [6] is an automata-based formalism to specify both control- and data-oriented properties. *ppDATE*s are basically transition systems with states and transitions between states. Transitions are labelled by a trigger (tr), a condition (c), and an action (a). Together, the label is written $tr \mid c \mapsto a$. A transition is *enabled* to be taken whenever its trigger is active and its condition holds. A trigger is activated by the occurrence of either a visible system event such as the invocation or termination of a method execution, or an action event generated by certain actions labelling other transitions. If a transition is taken, we will say that it *fires*. The conditions may depend on the values of *system variables* (i.e., variables of the system under scrutiny) and the values of *ppDATE variables*. The latter can be modified via actions in the transitions. *ppDATE* states represent the status of an *observer* of a system (rather then, directly, the status of a system itself). Note that each state essentially represents the set of observed system traces leading to that state. The language also offers parallelism on the specification side, in the sense that different *ppDATE*s



start
q :
(*i*) {cups < limit} brew() {cups == \old(cups)+1}
(*ii*) {true} cleanF() {cups == 0}
t_2 : brew$^{\uparrow}$ | *true* ↦ *skip*
t_1 : brew$^{\downarrow}$ | cups < limit ↦ *skip*
q' :
(*iii*) {cups < limit} brew() {cups == \old(cups)}
(*iv*) {true} cleanF() {cups == \old(cups)}
t_4 : cleanF$^{\downarrow}$ | *true* ↦ *skip*
t_3 : brew$^{\downarrow}$ | *true* ↦ *skip*
bad

Fig. 2. A *ppDATE* controlling the brew of coffee

run in parallel, possibly communicating which each other through events, and possibly creating new *ppDATE*s on demand. This parallelism allows for a strong separation of concerns in the specification.

In addition to the above, a particular feature of the *ppDATE* is that states may be tagged with any number of Hoare triples, to specify the computation of a method in a history-context sensitive way. For instance, assume that a *ppDATE* state q is tagged with the Hoare triple $\{\pi\}foo\{\pi'\}$. This means that, if *foo* is invoked after a system trace which led the observer to q, and if furthermore π holds at the time of the invocation, then π' should be satisfied upon termination of this execution of *foo*. This allows for data-centric specification of individual methods' behaviour (Hoare triple), however in a control sensitive manner (state).

Compared to usual automata based (or temporal logic based) specification approaches, *ppDATE* is more expressive concerning the computation on data. Compared to data-centric pre/post-specification (like, e.g., JML), *ppDATE* can avoid the coding of some notion of status into additional data and additional constraints in the pre/postconditions.

In the following examples, we tag transitions with labels of the form $tr \mid c \mapsto a$, where tr is the trigger of the transition, c is the condition which has to hold when tr occurs for the transition to be taken, and a is an action to be executed upon taking the transition. In addition, $\mathtt{tr}^{\downarrow}$ means that method associated to the trigger tr has just been called, and $\mathtt{tr}^{\uparrow}$ means that method associated to the trigger tr has terminated its execution.

Example 1. Let us consider a *coffee machine system* where after a certain amount of coffee cups are brewed, its filters have to be cleaned. If the limit of coffee cups is reached, the machine should not be able to brew any more coffee. In addition, while the coffee machine is active (a coffee cup is being brewed), it is not possible to start brewing another coffee, or to clean the filters.

Fig. 2 illustrates a *ppDATE* describing this part of the system. In other words, whenever the coffee machine is not active, i.e., the machine is not brewing a cup of coffee, and the method `brew` starts the coffee brewing process, then it is not possible either to execute this method again, or to execute the method `cleanF` (which initialises the task of cleaning the filter), until the initialised brewing process finishes.

The previous property can be interpreted as follows: initially being in state q, state which represents that the coffee machine is not active, whenever method `brew` is invoked and it is possible to brew a cup of coffee (i.e., the limit of coffee cups was not reached yet), then transition t_1 shifts the *ppDATE* from state q to state q'. While in q', state which represents that the coffee machine is active, if either method `brew` or method `cleanF` are invoked, then transitions t_3 or transition t_4 shift the *ppDATE* to state *bad*, respectively. This indicates that the property was violated. On the contrary, if method `brew` terminates its execution, then transition t_2 shifts the *ppDATE* from state q' to state q. Note that the names used on the transitions, e.g. t_1, t_2, etc, are not part of the specification language. They are included to simplify the description of how the *ppDATE* works.

In addition to this, the Hoare triples in state q ensure the properties: (i) if the amount of brewed coffee cups has not reached its limit yet, then a coffee cup is brewed; (ii) cleaning the filters sets the amount of brewed coffee cups to 0. Property (i) has to be verified if, while the *ppDATE* is on state q, the method `brew` is executed and its precondition holds. A similar situation stands for the property (ii) with respect to the method `cleanF`. Regarding state q', the Hoare triples in this state ensure the properties: (iii) no coffee cups are brewed; (iv) filters are not cleaned. Property (iii) and (iv) are verified if either method `brew` and method `cleanF` are executed, and their preconditions hold, respectively. Here, remember that this state represents that the coffee machine is active. Thus, if it occurs that either the method `brew` or the method `cleanF` are executed while the *ppDATE* is on this state, then, as this would move the *ppDATE* to state `bad`, one would expect the value of the variable `cup` to remain unchanged. This is precisely what is verified when either property (iii) or (iv) are analysed.

Note that none of the Hoare triples makes reference to the state of the coffee machine, i.e. there is no information about whether the machine is active or not. This is due to fact that the state of the machine is implicitly defined by the states of the *ppDATE*. If the *ppDATE* is in state q, the coffee machine is not active. However, if it is in state q', then the machine is active. Therefore, it is possible to assume that on each state the Hoare triples are *context dependent* and thus contain such information. This is the reason why, we can describe properties with the same precondition, but with different postconditions depending on the state of the *ppDATE* in which they are placed.

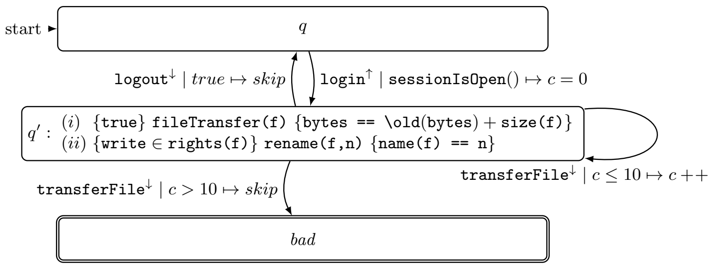

start
q
logout↓ | true ↦ skip
login↑ | sessionIsOpen() ↦ c = 0
q′ : (i) {true} fileTransfer(f) {bytes == \old(bytes) + size(f)}
(ii) {write ∈ rights(f)} rename(f,n) {name(f) == n}
transferFile↓ | c ≤ 10 ↦ c++
transferFile↓ | c > 10 ↦ skip
bad

Fig. 3. A *ppDATE* limiting file transfers

□

Example 2. In this example let us consider a *file system* where only 10 file transfers can be performed between a log in and log out of a user.

Fig. 3 illustrates a *ppDATE* describing part of the behaviour of this system. This *ppDATE* ensures the property: *no more than 10 file transfers take place in a single login session.* In other words, once a user logs in the system (`login`), she can only perform 10 file transfers (`transferFile`) before logging out (`logout`). This fact is tracked using the *ppDATE* variable c. This variable keeps count of the number of files transferred in a single session. Whenever a user logs in, the *ppDATE* moves to state q' and c is set to 0 (zero). While in q', this variable is increased by one every time a file transfer is performed. If at some point the user transfers a file but the value of c is bigger than 10, then the *ppDATE* moves to state *bad*, i.e., the property was violated.

In addition to this, the Hoare triples in state q' ensure the properties: (i) the number of bytes transferred increases when a file transfer is done; (ii) renaming a file works as expected if the user has the sufficient rights. □

4 The StaRVOOrS Framework

The StaRVOOrS framework (STAtic and Runtime Verification of Object-ORiented Software), originally proposed in [8], combines the use of the deductive source code verifier KeY [5] with that of the runtime monitoring tool Larva [20], to analyse and monitor systems with respect to a *ppDATE* specification. Note that the definition of the specification language *ppDATE*, which enables the effective combination of the results from the two verification approaches, is a major contribution of StaRVOOrS. *ppDATE* allows our framework to naturally address the intrinsic differences between the verification tools — whereas one typically verifies data-centric properties in deductive verifiers like KeY, one typically focuses on control-flow properties using runtime verifiers like Larva.

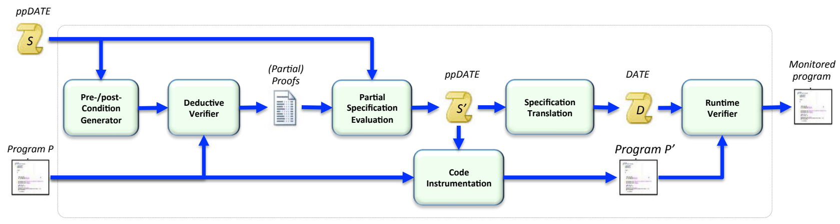


Fig. 4. High-level description of the StaRVOOrS framework workflow

The abstract workflow of the use of StaRVOOrS is given in Fig. 4. This workflow is applied fully automatically in four consecutive stages: *Deductive Verification*, *Specification Refinement*, *Translation and Instrumentation*, and *Monitor Generation*.

In the ***Deductive Verification*** stage, given a Java program P and a *ppDATE* specification S, the module Pre/post-Condition Generator transforms all the Hoare triples—assigned to the various states of S—into JML contracts, which are textually added to P as annotations of the respective methods. In this step, the association of pre/postcondition pairs to *ppDATE* states in S is lost, which is intentional and natural. Note that each *ppDATE* state represents the set of event histories leading to that state. The deductive verifier, however, offers analysis of the effect of methods *in terms of system data*, and has no notion of the history of events preceding a method call.[3] Once all JML contracts are generated, the Deductive Verifier module uses KeY in an attempt to statically verify each of them. The result is either a complete proof, or a partial proof where some branches are closed and others are not (see Sec. 2.1), or an entirely open proof, where no branches are closed. In our setting, partial proofs are the most common case. One reason is that we use KeY only fully automatically, not employing its interactive features. Also, we do not assume users to provide loop invariants, or similar annotations which support the prover. Finally, KeY has no knowledge of the *context* (*ppDATE* state) in which the Hoare triple at hand should hold. To illustrate this point, consider the Hoare triples (i) and (iii) from our (deliberately primitive) example in Fig. 2. The implementation of `brew()` is given by:

```
public void brew() {
    if (!active && cups < limit)
```

[3] There exist approaches to deductive verification which are history-aware, including a KeY version for the compositional verification of distributed systems [7]. These approaches are however much more heavyweight, both in terms of specification as well as verification, than what we are aiming at in this work. The same holds for approaches based on refinement.

```
        cups++;
    }
```

KeY will produce partial proofs, only, for (i) and (iii), because the specification does not provide information on how q and q' relate to the field `active`. In general, the missing information can be an arbitrary condition on the system state, more than just a Boolean as is the case here.

In the ***Specification Refinement*** stage,[4] the Partial Specification Evaluation module evaluates the results produced by KeY in order to refine S. This refinement is performed in two steps. In the first step, all fully verified Hoare triples are deleted, resulting in a *ppDATE* S'. Any Hoare triple related to a contract which is not fully verified by KeY is left in the states of S' to be verified at runtime. In the second step, S' is refined into a *ppDATE* S'' by strengthening the preconditions of those Hoare triples in S' which were *partially verified* by KeY. For that, the partial KeY proofs are analysed, to extract branch conditions corresponding to the closed branches of the proof. In the example in Sec. 2.1, that 'closed branch condition' is $\mathsf{y\%2} = \mathsf{0}$ in sequent (4). Note again that the branch condition is a condition *on the prestate* of the code being verified. Let us abbreviate the 'closed branch(es) condition' as *cbc* for now. A Hoare triple $\{\pi\}\mathit{foo}\{\pi'\}$ that was partially verified by KeY is clearly equivalent to having two Hoare triples $\{\pi \wedge \mathit{cbc}\}\mathit{foo}\{\pi'\}$ and $\{\pi \wedge \neg\mathit{cbc}\}\mathit{foo}\{\pi'\}$. However, as we know that the first one is valid (by the proof replay argument from Sec. 2.1), only the second one needs to be checked at runtime. For this reason, every Hoare triple $\{\pi\}\mathit{foo}\{\pi'\}$ in S' that was partially verified by KeY is replaced by $\{\pi \wedge \neg\mathit{cbc}\}\mathit{foo}\{\pi'\}$, resulting in S''. At runtime, checking such an optimised Hoare triple is trivial whenever π is false or *cbc* is true, as the postcondition does not need to be checked then. For instance, analysis of the partial proof of Hoare triple (i) in Fig. 2 will result in the closed branch condition ¬`active`. Therefore, (i) is replaced by $\{$`cups` $<$ `limit` $\wedge$ `active`$\}$ `brew()` $\{$`cups == \old(cups)+1`$\}$ (we simplified away double negation). Note that, in cases where the history context, i.e., *ppDATE* state, is the *only* information that was missing to close a partial proof, *cbc* actually represents a refinement of the according *ppDATE* state to a condition on internal system data, which will always be true when *foo* is called in that state. We can remark already here that this is the phenomenon which made the monitoring speedup particularly dramatic in the Mondex case study, see Sec. 10.

In the ***Translation and Instrumentation*** stage, the Specification Translation module translates S'' into an equivalent specification in *DATE* format (D), which can be used by the runtime verifier Larva (see the next stage). The most significant change of this translation is that the Hoare triples are translated away, using notions native to *DATE* (see Sec. 7.2).

[4] For readability, we use $\wedge$ and $\neg$ in this paragraph, instead of the *ppDATE* syntax `&&` and `!`.

This change also requires to instrument P, through the Code Instrumentation module, in order to (i) distinguish between different executions of the same code unit, and to (ii) evaluate Hoare triples in the states of S'' at runtime. Regarding (i), method declarations get a new argument which is used as a counter for invocations of this method. Regarding (ii), not every condition in a pre/postcondition of a Hoare triple can be directly written as a Java Boolean Expression, e.g., quantified expressions. Thus, methods which operationalise the evaluation of those conditions are added to P.

Finally, in the ***Monitor Generation*** stage, the instrumented version of P (P') and the *DATE* specification D are used by the Runtime Verifier module to generate a monitor M. For this, Larva generates M from D by using aspect-oriented programming techniques to capture relevant system events. Such events allow to link P' with M. Later, once deployed, M and P' are executed together. If M identifies any violation at runtime, it will report an error trace for further analysis.

5 Formal Definition of *ppDATE*s

5.1 Notation

We will use the following notation to write quantified formulae, based on the notation used by Gries [27].

$$\begin{gathered} \forall\ x \cdot R(x) \cdot B(x) \\ \exists\ x \cdot R(x) \cdot B(x) \end{gathered}$$

These formulae mean "for all x satisfying R, B is fulfilled" and "there exists x satisfying R for which B is fulfilled", respectively. Both R and B are formulae potentially containing x as a free variable. We will refer to R and B as the *range* and *body* of the quantified formula, respectively. This notation relates to standard (un-ranged) quantified formulae in the following way:

$$\begin{gathered} \forall\ x \cdot R(x) \cdot B(x) \equiv \forall\ x \cdot (R(x) \to B(x)) \\ \exists\ x \cdot R(x) \cdot B(x) \equiv \exists\ x \cdot (R(x) \land B(x)) \end{gathered}$$

5.2 *ppDATE*

In this section we formally define the notion of *ppDATE* previously introduced in Sec. 3. In order to do so, we first introduce formal definitions for triggers, conditions and actions.

Definition 1. *Given a set of method names Σ, the syntactic category of triggers is defined as follows:*

$$
\begin{aligned}
\textit{trigger} ::= \;& \textit{systemtrigger}\\
\mid\;& \mathsf{\textit{actevent?}}
\end{aligned}
$$

$$\textit{systemtrigger} ::= \mathsf{\textit{methodname}}^{\downarrow} \mid \mathsf{\textit{methodname}}^{\uparrow}$$

where $\mathsf{\textit{methodname}} \in \Sigma$. □

In the previous definition, *systemtrigger* matches a visible system event, such as the point of entry into a method or the termination of a method execution. Given a method name $\sigma \in \Sigma$, $\sigma^{\downarrow}$ represents entering method σ and $\sigma^{\uparrow}$ represents the termination of the execution of σ.

In addition, actevent represents an event generated by the execution of an action in a transition of a *ppDATE*, which we will call *action events*. This kind of events can only be generated by bang ("!") actions (see Def. 2). An action $h!$ generates the action event h, which in the next step can activate the trigger $h?$. This way, action events enable communication among *ppDATE*s, where $h!$ and $h?$ mean sending and receiving a message, respectively.

As we have mentioned before, whenever a transition is fired an action can be executed. The following shows the definition of actions.

Definition 2. *Actions are syntactically defined as follows:*

$$
\begin{aligned}
\textit{action} ::= \;& \mathsf{\textit{skip}}\\
\mid\;& v = e\\
\mid\;& \mathsf{\textit{actevent!}}\\
\mid\;& \mathsf{\textit{create}}(\textit{template}, \overline{\textit{args}})\\
\mid\;& \textit{action} \;;\; \textit{action}\\
\mid\;& \mathsf{\textit{if}}\ \textit{cond}_{Sys \cup V}\ \mathsf{\textit{then}}\ \textit{action}\\
\mid\;& \mathsf{\textit{Program}}
\end{aligned}
$$

□

skip is the effect-less action. The '=' is an assignment operator, v is a *ppDATE* variable and e is a (side-effect free) expression that may depend on system variables and *ppDATE* variables; actevent! represents the generation of action event actevent; create represents the creation of a *ppDATE*, where *template* is a *ppDATE* template to be instantiated (see Def. 8), and $\overline{\textit{args}}$ are the values which the formal parameters of *template* are instantiated with; the ';' is the sequence operator for actions; if-then is a conditional whose branching condition depends on the valuations of system variables (*Sys*) and *ppDATE* variables (V); and Program represents a *side-effect free* program (see Def. 3), i.e., it is restricted to not have any effect on the system which could in turn be observed by the (*ppDATE* generated) monitor. For instance, a Program could perform logging of system/monitor behaviour. More powerful Programs, which would for instance allow error recovery, are relevant, but left for future work.

Definition 3. *A* side-effect free program *has the properties that*

- *its execution always terminates,*
- *the method calls on its body do not generate any observable system event,*
- *it does not interfere with the system under scrutiny, i.e., it does not modify the values of system variables.*

□

Boolean expressions are used in different contexts: (i) conditions (c) of transitions; (ii) conditions of `if-then` actions, and (iii) pre- and postconditions (π, π') in Hoare triples. As a syntactic category for such Boolean expressions, we chose *Boolean JML expressions*. They extend *Boolean Java expressions*, and thereby allow Java methods as sub-expressions (like in '`m.get(k) == o`'). Additional features of *Boolean JML expressions* include universal and existential quantification, which are frequently used in Hoare triples, the ability to refer in a postcondition to a) the return value (with '`\result`'), and b) the preexecution value of an expression (like in '`x == \old(x + y)`').

Definition 4. Boolean JML expressions *(BJMLE) are recursively defined as follows:*

- *any side-effect free Boolean* Java *expression is a BJMLE,*
- *if* `a` *and* `b` *are BJMLEs, and* `x` *is a variable of type* `t`*, the following expressions are BJMLEs:*
  - `!a`, `a&&b`, *and* `a||b`
  - `a ==> b` *("*`a` *implies* `b`*")*
  - `a <==> b` *("*`a` *is equivalent to* `b`*")*
  - `(\forall t x; a)`
    ("for all `x` *of type* `t`*,* `a` *holds")*
  - `(\exists t x; a)`
    ("there exists `x` *of type* `t` *such that* `a`*")*
  - `(\forall t x; a; b)`
    ("for all `x` *of type* `t` *fulfilling* `a`*,* `b` *holds")*
  - `(\exists t x; a; b)`
    ("there exists an `x` *of type* `t` *fulfilling* `a`*,*
    such that `b`*")*
- *replacing any sub-expression* `e` *in a BJMLE with* `\old(e)` *gives a BJMLE,*
- *replacing any sub-expression in a BJMLE with* `\result` *gives a BJMLE, (well-typedness is context dependent, see Def.5)*

□

We do not give a formal definition of the semantics of BJMLE here, just the following comments. The meaning of negation, conjunction, disjunction, implication, and equivalence are standard. The same is true for the first two forms of quantification. Concerning the other two forms, "... `a; b)`", they relate to standard quantification in exactly the same way as was explained in Sec. 5.1. (The only difference is that there we discussed meta-level notation, whereas BJMLE is part of *ppDATE*.) The constructs `\old` and `\result` are only allowed in postconditions of Hoare-triples (i.e., in π'). `\result` refers to the return value of a (non-`void`) method. `\old` allows to evaluate sub-expressions not in the post-state (which is the default), but in the prestate of a method's execution. For instance, '`x == \old(x + y)`' in a postcondition of method `m` says that the difference between the values of `x` before and after the execution of `m` is the value which `y` had *before* `m`'s execution.

In order to allow or disallow `\old` and `\result`, in the following, we provide one syntactic category for postconditions, and one for all other conditions.

Definition 5. *The syntactic category of postconditions over variables in Var, $postcond_{Var}$, is given by Boolean JML expressions over Var. (Well-typedness of postconditions is context dependent, assuming that `\result` has the same type as the specified method.) The syntactic category $cond_{Var}$ is given by Boolean JML expressions over Var containing neither `\result` nor `\old`.* □

Now we can formally define *ppDATE*. As a *ppDATE* describes properties about a particular system, we assume that every time we make reference to the set of system variables, these variables belong to the system under scrutiny.

Definition 6. *Given a set of system variables Sys and a set of ppDATE variables V, a ppDATE m is a tuple (Q, t, B, q_0, Π) such that:*

- *Q is the finite set of states.*
- *t is the transition relation among states in Q, where each transition is tagged with (i) a trigger; (ii) a condition; (iii) an action which may change the valuation of ppDATE variables: $t \subseteq Q \times trigger \times cond_{Sys \cup V} \times action \times Q$.*
- *$B \subseteq Q$ is the set of bad states.*
- *$q_0 \in Q$ is the initial state.*
- *Π is a function which tags each state of m with Hoare triples for particular method names in Σ: $\Pi \in Q \longrightarrow \mathcal{P}(cond_{Sys} \times \Sigma \times postcond_{Sys})$.* □

We will write $q \xrightarrow{tr|c\mapsto a}_m q'$ to mean that, given a *ppDATE* m whose transition relation is t, $(q, tr, c, a, q') \in t$. The subscript m is omitted if it is clear from the context. In addition, we will use the usual Hoare triple notation $\{\pi\}\,\sigma\,\{\pi'\} \in \Pi(q)$ to denote $(\pi, \sigma, \pi') \in \Pi(q)$.

Example 3. Consider once again, the *ppDATE* shown in Fig. 3. It can be formalised as follows: $m = (Q, t, B, q_0, \Pi)$, where,

- $Q = \{q, q', \mathit{bad}\}$,
- $V = \{c\}$,
- $\Sigma = \{\texttt{fileTransfer}, \texttt{login}, \texttt{logout}\}$,
- $B = \{\mathit{bad}\}$,
- $q_0 = q$.

Furthermore, the transition relation t consists of four elements, including: $q' \xrightarrow{\texttt{fileTransfer}^{\downarrow} | c \leq 10 \mapsto c{+}{+}} q'$ and $q' \xrightarrow{\texttt{fileTransfer}^{\downarrow} | c > 10 \mapsto \texttt{skip}} \mathit{bad}$. In addition, relation Π is defined as follows:

$$\begin{aligned}
\Pi(q) &= \{\ \{\texttt{true}\}\ \texttt{fileTransfer(f)}\ \{\texttt{bytes} == \texttt{\textbackslash old(bytes)}\}\ \}\\
\Pi(q') &= \{\ \{\texttt{true}\}\ \texttt{fileTransfer(f)}\ \{\texttt{bytes} == \texttt{\textbackslash old(bytes)} + \texttt{size(f)}\},\\
&\qquad \{\texttt{write} \in \texttt{rights(f)}\}\ \texttt{rename(f,n)}\ \{\texttt{name(f)} == n\}\ \}
\end{aligned}$$

□

In addition to *ppDATE*s which exist up-front, and 'run' from the beginning of a system's execution, new *ppDATE*s can be created by existing ones. For instance, one may want to create a separate 'observer' for each new user logged into a system. For that, one needs to be able to define parameterised *ppDATE*s, which we call *templates*, and allow *ppDATE*s to create new instantiations of templates. Given a *ppDATE* m, the creation of a new *ppDATE*, which will run in parallel to m, can be achieved by using action create on a transition of m. This action receives as arguments a *ppDATE* template describing the *ppDATE* to be created and a list of arguments to instantiate the quantified variables on the template. Below, we formally define *ppDATE* templates.

Definition 7. *ppDATE* templates of order n *are recursively defined as follows:*

- *The set of* ppDATE templates of order 0 *is exactly the set of ppDATEs.*
- *Assume C is a syntactic sub-category of* ppDATE *(Def. 6), i.e., a syntactic (sub-)category of Q, t, B, q_0, or Π, respectively. If m is a ppDATE template of order n, then $\lambda X{:}C.m'$ is a ppDATE* template of order $n+1$*, where m' is the result of replacing, in m, some (sub-)term trm of category C by X. We call X the template variable of $\lambda X{:}C.m'$.* □

In the above definition, a template of order $n + 1$ is defined by 'abstracting' over templates of order n, annotating the abstracted 'hole' X by the right category, such that template instantiation (see below) can be guaranteed to result in a well-typed *ppDATE*. When constructing a *ppDATE* template, the choice of *trm* in Def. 7 does not matter. Its only role is to carry well-typedness of *ppDATE*s over to *ppDATE* templates. Informally, the above definition says that, within $\lambda X{:}C.m'$, the X can appear anywhere in m' where a term of category C is expected.

We will refer to *ppDATE* templates without referring to an order to mean templates that are not of order greater than 0. Formally:

Definition 8. *The set of ppDATE* templates T_{ppd}*, is defined as the union of ppDATE templates of order* $n \geq 1$*.*

If $\overline{X}$ is a vector of template variables $X_1, \ldots, X_n$ and $\overline{C}$ is a vector of syntactic categories $C_1, \ldots, C_n$, then we can write $\lambda\overline{X}{:}\overline{C}.m$ to mean $\lambda X_1{:}C_1 \ldots \lambda X_n{:}C_n.m$.

Finally, we define what it means to instantiate a *ppDATE* template:

Definition 9. *Given a term trm of syntactic category C, the* instantiation of a *ppDATE* template with term *trm, denoted inst*(m, trm)*, is defined by:*

$$inst(\lambda X{:}C.m,\ trm) = m[X/trm]$$

where $m[X/trm]$ *denotes the result of substituting all occurrences of* X *in* m *by trm.*

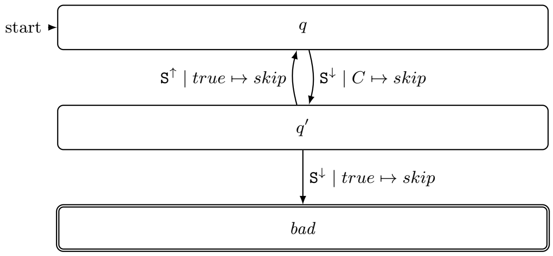
We can expand template instantiation to multiple arguments in the following way. Given $n \geq 2$, assume $\overline{X} = X_1, \ldots, X_n$, and $\overline{C} = C_1, \ldots, C_n$, and $\overline{trm} = trm_1, \ldots, trm_n$ (with $trm_i \in C_i$). We extend the instantiation function *inst* to an arbitrary number of arguments in the following way:

$$\begin{array}{l} inst(\lambda\overline{X}{:}\overline{C}.m,\ \overline{trm}) \\ = \textit{(by syntactic convention)} \\ inst(\lambda X_1{:}C_1 \ldots \lambda X_n{:}C_n.m,\ trm_1, \ldots, trm_n) \\ \overset{df}{=} \\ inst(inst(\lambda X_1{:}C_1 \ldots \lambda X_n{:}C_n.m,\ trm_1),\ trm_2, \ldots, trm_n) \end{array}$$

Example 4. Fig.5 illustrates a *ppDATE* template, based on the *ppDATE* depicted in Fig. 2. Let us call it *one-at-a-time*. This template has two parameters: C, which represents a condition, and S, which represents a method name. Then, by executing the action `create`(*one-at-a-time*, $cups < limit$, $\texttt{brew}^{\downarrow}$), it would instantiate the *ppDATE* depicted in Fig.6, i.e., C is instantiated with $cups < limit$ and S is instantiated with `brew`. This *ppDATE* specifies the property: *it is not possible to brew one more coffee cup until the brewing process is done.* □

In the rest of this work we will only consider the use of deterministic *ppDATE*s. Formally:

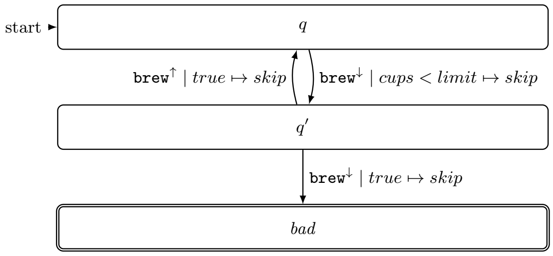
$\mathit{one\text{-}at\text{-}a\text{-}time} = \lambda\ C, S : \mathit{cond}, \mathit{trigger}.$



start
q
$\mathtt{S}^{\uparrow} \mid \mathit{true} \mapsto \mathit{skip}$
$\mathtt{S}^{\downarrow} \mid C \mapsto \mathit{skip}$
q'
$\mathtt{S}^{\downarrow} \mid \mathit{true} \mapsto \mathit{skip}$
bad

Fig. 5. *ppDATE* template example.

$\mathit{inst}(\mathit{one\text{-}at\text{-}a\text{-}time}, \mathit{cups} < \mathit{limit}, \mathtt{brew}) =$



start
q
$\mathtt{brew}^{\uparrow} \mid \mathit{true} \mapsto \mathit{skip}$
$\mathtt{brew}^{\downarrow} \mid \mathit{cups} < \mathit{limit} \mapsto \mathit{skip}$
q'
$\mathtt{brew}^{\downarrow} \mid \mathit{true} \mapsto \mathit{skip}$
bad

Fig. 6. *ppDATE* created using the template illustrated in Fig. 5.

Definition 10. *We say that a ppDATE m is* deterministic *if, for any two transitions of m with same trigger tr which go from a state q to a different state, their conditions are mutually exclusive:*

$$\begin{aligned} &\forall\ tr, c, c', a, a', q, q', q'' \cdot \\ &\quad q \xrightarrow{tr|c \mapsto a}_m q' \ \mathit{and}\ q \xrightarrow{tr|c' \mapsto a'}_m q'' \cdot \mathit{not}(c\ \mathit{and}\ c') \end{aligned}$$

□

In addition, although determinism on the Hoare triples' preconditions is not problematic in itself, we choose to extend the determinism condition to ensure that for any two Hoare triples in a single state over the same function have disjoint precondition so as to have a more effective monitoring algorithm of these triples: for any $\{\pi_1\}\,\sigma\,\{\pi_1'\}$ and $\{\pi_2\}\,\sigma\,\{\pi_2'\}$ in $\mathit{\Pi}(q)$, $\mathit{not}(\pi_1\ \mathit{and}\ \pi_2)$.

After having defined (individual) *ppDATE*s, we can now define a *network* of *ppDATE*s.

Definition 11. *Given a set of system variables Sys, a ppDATE* network *pn is represented with a tuple (M, V, ν_0, T_{ppd}):*

– *M is a set of ppDATEs. If $m \in M$, then we say that $m = (Q_m, t_m, B_m, q_{0m}, \Pi_m)$.*
– *V is a set of ppDATE variables.*
– *ν_0 is the initial valuation*[5] *of variables in V.*
– *T_{ppd} is a set of ppDATE templates.* □

Note that on a network, whenever a trigger is activated, several *ppDATE*s can have an enabled transition ready to be fired, i.e., a transition whose trigger is active and whose condition holds. Whenever this happens all these enabled transitions are fired in parallel. Also note that the set of *ppDATE* variables V is global to the network of *ppDATE*s, rather than local to individual *ppDATE*s. Thereby, V is effectively the 'shared memory' of the network.

Finally, we extend the notion of deterministic *ppDATE* to a *ppDATE* network.

Definition 12. *A ppDATE network $pn = (M, V, \nu_0, T_{ppd})$ is* deterministic *whenever every ppDATE in M is deterministic and every ppDATE which can be created when executing action* create *is deterministic.* □

6 *ppDATE* semantics

In this section we present the semantics of a network of *ppDATE*s by introducing *Structural Operational Semantics* (SOS) rules. These rules will show how a global configuration is shifted to a new one by considering events and system variables valuations in a system trace.

Informally, a global configuration (L, ν) (of a *ppDATE* network) consists of a set L of local configurations (one for each *ppDATE* in the set of *ppDATE*s of the network and one for each generated instance of a *ppDATE* template), and a valuation ν of the set of *ppDATE* variables V (associated to the *ppDATE* network). The local configurations store the current state, and record, for each ongoing method execution whose precondition was fulfilled at call time, the postcondition to be checked on exit.

Every time the system under scrutiny generates an event, e.g., by entering or leaving a method, every local configuration in L which has an enabled transition will, simultaneously, replace its current state value by the state indicated in the fired transition, and execute the action of this transition, also simultaneously. For instance, given a *ppDATE* m whose current state is q, and with a transition t_1 of the form $q \xrightarrow{tr|c \mapsto a}_m q'$, if a system event that activates trigger tr occurs and condition c holds, then t_1 is fired, state q is replaced by q' in the appropriate local configuration in L, and a is executed. Regarding the executed actions, if they contain

[5] A valuation is a mapping from variables to values of adequate types.

ppDATE variables assignments, those assignments change the valuation ν. In addition, if these executions generate action events, those events will be stored in a buffer.

Once all the previous enabled transitions are fired, every transition that become enabled by the events in the buffer will be fired as well. For instance, let us assume that action a in transition t_1 (only) generates the action event h, i.e., $a = h!$, and that a *ppDATE* m' running in parallel to m is in state q'', and has a transition t_2 of the form $q'' \xrightarrow{h?|true \mapsto a'}_{m'} q'''$. Then, whenever t_1 is fired, execution of $h!$ will add to the buffer an event which will enable t_2, due to the fact that trigger $h?$ is activated by h and its condition (trivially) holds. Therefore, after firing t_1, t_2 will be also fired.

Note that before firing all the transitions enabled by the events in the buffer, the buffer will be emptied. Therefore, the buffer only contains events generated by the recent action executions, and no events from previous ones. This procedure is repeated until no new action event is generated, i.e., the buffer is empty.

6.1 Events, Valuations, and Traces

ppDATE networks describe which system behaviours are allowed, and which are not. Here, we consider as behaviour basically a series of system events, where each event also comes with a 'snapshot' of the values of (visible) system variables, taken at the time where the event occurs. Formally, these snapshots are *valuations*, i.e., mappings from variables to values (of adequate types). Apart from the observed system, the *ppDATE* networks themselves may create new events.

An event may therefore either be a *system event* (i.e., an event generated by the system under scrutiny due to entering or leaving a method) or an *action event* (i.e., an event generated by the execution of an action ! in a *ppDATE* transition). Formally:

Definition 13. *Given a set of method names Σ, the syntactic category of events is defined as follows:*

$$\xi ::= systemevent \mid \mathsf{actevent}$$

$$systemevent ::= systemtrigger_{\mathbb{N}} \qquad \square$$

A *systemevent* consists of a *systemtrigger* which is indexed with a natural number representing *nth* execution of the method associated to the trigger. Such an index will be considered an identifier[6] unique to each execution of the method.

[6] These identifiers can be created automatically using techniques as those presented in [24] or through stack frame references.

We distinguish the set of system variables valuations Θ_{Sys}, with typical element θ, and the set of *ppDATE* variable valuations N, with typical element ν. We represent valuations both as functions and (functional) relations[7], i.e., sets of pairs. This means that the notation $\beta(v) = val$ is equivalent to the notation $(v, val) \in \beta$. The *union of valuations* is therefore a set union such that, for any two valuations β and β', $\beta \cup \beta' = \{(v, val) \mid (v, val) \in \beta \textit{ or } (v, val) \in \beta'\}$. In the presentation of examples, we limit the valuations to those variables which matter for the example at hand, for simplicity.

In our semantic rules, we will use union over valuations only when the domain of valuations, the variable sets, do not overlap, like for instance in $\theta \cup \nu$. Another operation on valuations is the *modification* of a valuation β at variable x by value val, written $\beta[x \leftarrow val]$. It is defined as:

$$\beta[x \leftarrow val](v) = \begin{cases} val & \text{iff } \; v = x \\ \beta(v) & \text{otherwise} \end{cases}$$

Given a set of variables S, a valuation β for S, and condition $c \in cond_S$, we will write $\beta \models c$ to denote that c is satisfied by β. This is however not sufficient for postconditions, as they can refer to two valuations, after and before ("`\old`") a method's execution. For that, $\models$ will be overloaded. Given a set of system variables *Sys*, valuations θ and θ' for it, and a postcondition $c \in postcond_{Sys}$, we will write $\theta, \theta' \models c$ to denote that c is satisfied by θ and θ'. When this is used, θ' will be the current valuation of *Sys* when exiting a certain method execution, whereas θ holds the valuation from before that method execution. We only sketch the definition of $\models$ here, as it mainly follows the standard of first-order logic semantics. We use the two semantic truth values T and F. For $c \in cond_S$, we define $\beta \models c$ *iff* $eval_\beta(c) = T$, where $eval_\beta$ is recursively defined over the structure of c in the way which is standard in first-order logic[8], with the base case $eval_\beta(\mathtt{x}) = \beta(\mathtt{x})$ for variables `x`. In case of postconditions $c \in postcond_{Sys}$, we define $\theta, \theta' \models c$ *iff* $eval_{\theta,\theta'}(c) = T$. The definition of $eval_{\theta,\theta'}$ is almost identical to the definition $eval_\beta$, with the base case $eval_{\theta,\theta'}(\mathtt{x}) = \theta'(\mathtt{x})$ for program variables `x`. The only case in the definition where the other parameter, the prevaluation θ, matters is the evaluation of `\old`-expressions: $eval_{\theta,\theta'}(\mathtt{\backslash old(}e\mathtt{)}) = eval_\theta(e)$. This means that, in postconditions, the post-valuation θ' acts as the default, however not inside `\old`-expressions, where instead the prevaluation θ counts.

A *system trace* is a sequence of tuples consisting of an *event* and 'system snapshot', i.e., a valuation of the system variables, taken at the time when that event occurs.

[7] A (binary) relation R is *functional* if $\{(x, y), (x, y')\} \subseteq R$ implies $y = y'$.

[8] To be precise, *eval* has one extra parameter, which is a *logical* variable assignment, needed to define the evaluation of quantified formulas. We omit that parameter, because it is unimportant for our discussion here.

Definition 14. *A* system trace *w is a sequence of tuples in $systemevent \times \Theta_{Sys}$, i.e. $w \in (systemevent \times \Theta_{Sys})^*$. The first component of these tuples consists of a system event and the second one is the valuation of the set of system variables Sys.* □

6.2 Configurations

Given a system trace w, each tuple in w will shift a *global configuration* of a *ppDATE* network to another. Global configurations are defined with the help of local ones, so we start there.

Definition 15. *Given a set of method names Σ, a* local configuration *is a tuple (m, q, ρ) where m is a ppDATE, $q \in Q_m$, and $\rho \subseteq \mathcal{P}(systemevent \times postcond_{Sys} \times \Theta_{Sys})$.* □

The tuple (m, q, ρ) is a configuration of *ppDATE* m — where q represents the current state, and ρ allows to monitor potential violations of Hoare triples. For that, ρ stores which exit event ($\in$ *systemevent*) should cause a checking of which postcondition ($\in$ *postcond*). The semantic rules described below (Sec. 6.4) will guarantee that only method exit events (of the form $\sigma_i^{\uparrow}$) will appear in ρ. During the processing of a trace, the appearance of $(\sigma_i^{\downarrow}, \theta)$ at the same time as the current state has a Hoare-triple with a fulfilled precondition, $\theta \models \pi$, the corresponding postcondition π' is associated with $\sigma_i^{\uparrow}$ in ρ, together with θ. Later, the appearance of $(\sigma_i^{\uparrow}, \theta')$ will cause a look-up of $(\sigma_i^{\uparrow}, \pi', \theta)$ in ρ, in order to check $\theta, \theta' \models \pi'$.

Example 5. Recall the *ppDATE* illustrated in Fig. 2, here called m. Its initial local configuration is $(m, q, \emptyset)$. Then, after firing transition t_1 whenever certain system event $\texttt{brew}_{id}^{\downarrow}$ (with $id \in \mathbb{N}$) occurs, assuming that the field *cups* is valuated to zero, the next local configuration is $(m, q', \{(\texttt{brew}_{id}^{\uparrow}, cups == \texttt{\textbackslash old}(cups) + 1, \{(cups, 0)\})\})$. □

Definition 16. *Given ppDATE network $pn = (M, V, \nu_0, T_{ppd})$, a* global configuration *for pn is a tuple (L, ν) such that:*

- *L is a set of local configurations. For each $m \in M$, there is exactly one q and one ρ, such that $(m, q, \rho) \in L$. For each $(m, q, \rho) \in L$, we have $q \in Q_m$ and either $m \in M$ or $m = inst(t, \overline{args})$, for some $t \in T_{ppd}$.*
- *ν is ppDATE variable valuation with domain V.* □

Before giving an example, we define the notion of *initial global configuration* for a *ppDATE* network.

Definition 17. *Given ppDATE network $pn = (M, V, \nu_0, T_{ppd})$ where $m \in M$ is defined as a tuple $(Q_m, t_m, B_m, q_{0m}, \Pi_m)$, the* initial global configuration *$C_{init}(pn)$ is defined as the tuple (L_0, ν_0), where $L_0 = \{(m, q_{0m}, \emptyset) \mid m \in M\}$ is the set of initial local configurations.* □

Example 6. Let us assume that a *ppDATE* network $pn = (\{m, m'\}, \{v\}, \{(v, 0)\}, \emptyset)$, such that $q_{0m'} \xrightarrow{tr|true \mapsto v=v+1}_{m'} q_{1m'}$. The initial global configuration for pn is $C_{init}(pn) = (L_0, \{(v, 0)\})$, where $L_0 = \{(m, q_{0m}, \emptyset), (m', q_{0m'}, \emptyset)\}$. Then, if the given transition is fired, the new global configuration is $(L', \{(v, 1)\})$, where $L' = \{(m, q_{0m}, \emptyset), (m', q_{1m'}, \emptyset)\}$. □

The action in Example 6, $v = v + 1$, does not generate any event. But in general, actions may generate action events. For storing those (and ultimately process them in the next step), we introduce the concept of *extended global configuration.*

Definition 18. *Given a network $pn = (M, V, \nu_0, T_{ppd})$, and a set of system variables Sys, an* extended global configuration *for pn is a tuple (L, ν, E, θ) such that:*

- *(L, ν) is a global configuration for pn,*
- *$E \subseteq \mathcal{P}(\xi)$ is a set of events,*
- *$\theta \in \Theta_{Sys}$ is a system variables valuation.* □

E contains the events to be processed in the next (small) step. In the operational semantics to be described below, E will either be a singleton set containing a system event, or a set of action events generated by the executions of actions in the latest transition.

Example 7. Let us assume a *ppDATE* network $pn = (\{m, m'\}, \{v\}, \{(v, 0)\}, \emptyset)$, such that $q_1 \xrightarrow{\mathtt{foo}^{\downarrow}|true \mapsto h!}_{m} q_2$, $q_1' \xrightarrow{h?|true \mapsto v=v+1}_{m'} q_2'$, $\Pi_m(q_1) = \{\{\pi\}\mathtt{foo}\{\pi'\}\}$, and that q_1 and q_1' are the initial states of m and m', respectively. In addition, let us assume that $C_1 = (L_1, \{(v, 0)\}, \{\mathtt{foo}^{\downarrow}_{id}\}, \emptyset)$ is an extended global configuration for pn (for some index $id \in \mathbb{N}$), where $L_1 = \{(m, q_1, \emptyset), (m', q_1', \emptyset)\}$. Then, when the given transition of m is fired, given that π holds and the current system variables valuation is θ, the next extended global configuration for pn is $C_2 = (L_2, \{(v, 0)\}, \{h\}, \emptyset)$, where $L_2 = \{(m, q_2, \{(\mathtt{foo}^{\uparrow}_{id}, \pi', \theta)\}), (m', q_1', \emptyset)\}$. After that, event h in C_1 triggers the given transition of m', leading to the extended global configuration $C_3 = (L_3, \{(v, 1)\}, \emptyset, \emptyset)$, where $L_3 = \{(m, q_2, \{(\mathtt{foo}^{\uparrow}_{id}, \pi', \theta)\}), (m', q_2', \emptyset)\}$. □

The Structural Operational Semantics given in Sec. 6.4 formalises such behaviour.

6.3 Semantics of Actions

When assigning meaning to actions, there are two levels to consider. One is the level of the local actions, executed when an individual *ppDATE* takes a transition. The semantics of those is sequential, as defined below.

On top of the assignments changing the *ppDATE* variable valuation, the local actions may generate events, and create new instances of *ppDATE* templates. Both the events and new *ppDATE*s are simply accumulated, in sets of events and *ppDATE*s, respectively.

The other level is parallel actions, where we compose simultaneous actions of transitions taken in parallel by different *ppDATE*s. Here, we need to devote special care to excluding conflicting writes to, as well as race conditions between reads and writes from/to, the same variable. Also, we need to make sure that if one *ppDATE* writes to x, while the others do not, that the parallel composition keeps the effect of the first. All this makes it necessary to keep track of what is written to, and read from, already in the local actions, prior to parallel composition. However, the treatment of the local effects and newly created *ppDATE*s is simper. We just take the union of those when doing the parallel composition.

Definition 19. *For each action $a \in$ action, its* meaning $[\![a]\!]_{\theta,\nu}$ *(relative to system/ppDATE variable valuations θ and ν) is given by a tuple (ν', W, R, E, New), where:*

- $\nu' \in N$ *is a ppDATE variable valuation computed (locally) in a,*
- $W \subseteq V$ *is a set of ppDATE variables written to in a,*
- $R \subseteq V$ *is a set of ppDATE variables read from in a,*
- $E \subseteq$ actevent *is a set of action events generated in a,*
- $New \subseteq$ *ppDATE is a set of ppDATEs newly created in a.*

Given that pvars returns the ppDATE variables appearing in its argument(s), $[\![a]\!]_{\theta,\nu} = (\nu', W, R, E, New)$ is defined as follows

$$
\begin{aligned}
[\![\mathsf{skip}]\!]_{\theta,\nu} &= (\nu, \emptyset, \emptyset, \emptyset, \emptyset)\\
[\![v = e]\!]_{\theta,\nu} &= (\nu[v \leftarrow eval_{\theta\cup\nu}(e)], \{v\}, pvars(e), \emptyset, \emptyset)\\
[\![h!]\!]_{\theta,\nu} &= (\nu, \emptyset, \emptyset, \{h\}, \emptyset)\\
[\![\mathsf{create}(t, \overline{args})]\!]_{\theta,\nu} &= (\nu, \emptyset, pvars(\overline{args}), \emptyset, inst(t, \overline{args}))\\
[\![a_1 \ ; \ a_2]\!]_{\theta,\nu} &= \begin{cases} (\nu_2, W_1 \cup W_2, R_1 \cup R_2, E_1 \cup E_2, New_1 \cup New_2) \\ \quad \textit{where} \\ \quad [\![a_1]\!]_{\theta,\nu} = (\nu_1, W_1, R_1, E_1, New_1) \\ \quad \textit{and} \\ \quad [\![a_2]\!]_{\theta,\nu_1} = (\nu_2, W_2, R_2, E_2, New_2) \end{cases}\\
[\![\mathsf{if}\ c\ \mathsf{then}\ a]\!]_{\theta,\nu} &= \begin{cases} (\nu', W, R \cup pvars(c), E, New) \\ \qquad \textit{if } \theta \cup \nu \models c \textit{ and } [\![a]\!]_{\theta,\nu} = (\nu', W, R, E, New) \\ (\nu, \emptyset, pvars(c), \emptyset, \emptyset) \\ \qquad \textit{otherwise} \end{cases}\\
[\![prog]\!]_{\theta,\nu} &= [\![\mathsf{skip}]\!]_{\theta,\nu}
\end{aligned}
$$

□

Following the definition of actions (Def. 2), the *prog* in the last line above is a side-effect free program, i.e., it has no effect which could be noticed in the current formalism, which is why we can simulate it with skip. *prog* will have purposes orthogonal to our formalisation, like logging.

We are now in the position to define the parallel composition of actions. Imagine we have a configuration with 5 parallel *ppDATE*s, 3 of which have currently enabled transitions, with actions a_1, a_2, and a_3, respectively. Assume moreover that the current *ppDATE* variable valuation is ν. The parallel composition of the meaning of a_1, a_2, and a_3, i.e., of $[\![a_1]\!]$, $[\![a_2]\!]$, and $[\![a_3]\!]$, is performed by $\mathit{mergeParalActs}_\nu(\{[\![a_1]\!], [\![a_2]\!], [\![a_3]\!]\}) = (\nu', E', \mathit{New}')$. The function *mergeParalActs* takes a set of semantic actions as input, and computes a resulting valuation ν', a resulting set of events E', and a resulting set of newly generated *ppDATE*s, New'. The sets E' and New' will simply be the union of the corresponding sets from $[\![a_1]\!]$, $[\![a_2]\!]$, and $[\![a_3]\!]$. But the resulting valuation is slightly more involved. Actions may conflict (e.g., we write to the same variable in different actions), or have race conditions (i.e., we read from a variable and write to it in different actions). In those cases, we leave the result of *mergeParalActs* deliberately *undefined.* In all other cases, the different effects of the actions are merged. The index of the merging function, ν, serves as a fall back for those variables which have not been written to. In particular, the $\nu' = \nu$ in case the set of actions to be merged is empty.

These explanations are formalised in the following function, merging a set of action meanings (Def. 19):

Definition 20.
$$\mathit{mergeParalActs}_\nu(\{(\nu_1, W_1, R_1, E_1, \mathit{New}_1), \ldots, (\nu_n, W_n, R_n, E_n, \mathit{New}_n)\})$$
$$= \begin{cases} \mathit{undefined} \\ \quad \mathit{if}\ \exists\ i, j \cdot (i, j \in [1, .., n]\ \mathit{and}\ i \neq j) \cdot (W_i \cap W_j \neq \emptyset\ \mathit{or}\ W_i \cap R_j \neq \emptyset) \\ (\nu', E', \mathit{New}')\ \mathit{otherwise,\ where} \\ \quad E' = \bigcup_{i=1}^n E_i, \quad \mathit{New}' = \bigcup_{i=1}^n \mathit{New}_i,\ \nu'(v) = \begin{cases} \nu_i(v)\ \mathit{if}\ v \in W_i \\ \nu(v)\ \mathit{if}\ v \notin \bigcup_{i=1}^n W_i \end{cases} \end{cases}$$
□

Note that, if there are no actions to merge, we have $\mathit{mergeParalActs}_\nu(\emptyset) = (\nu, \emptyset, \emptyset)$.

6.4 Structural Operational Semantics

In this section we give structural operational semantics rules (SOS) for *ppDATE*s. These rules will have the following generic form:

$$\mathit{name}\frac{\begin{array}{c} H_1 \\ \cdots \\ H_n \end{array}}{\mathit{Goal}}$$

where *name* is a label used to identify the rule, *Goal* is the property enforced by the rule and the premises $H_1, \cdots, H_n$ are assumptions over the values of the *Goal*.

Predicate Definitions In the semantic definitions given below, we use the following predicates, as abbreviations.
activatedBy Given a (transition) trigger tr and an event e, predicate $\mathit{activatedBy}(tr, e)$ holds if tr and e match, in the following way:

$$\mathit{activatedBy}(tr, e) \overset{df}{=} \begin{cases} \exists\, i \cdot i \in \mathbb{N} \cdot e = tr_i & \text{iff } e \in \mathit{systemevent} \\ tr = e? & \text{iff } e \in \mathsf{actevent} \end{cases}$$

For instance, the trigger $\sigma^{\downarrow}$ is activated by the *systemevent* $\sigma^{\downarrow}_{3}$, and the trigger $h?$ is activated by $\mathsf{actevent}$ h (generated before by the execution of action $h!$).

$\square$

nextState Given a local configuration (m, q, ρ), a state q', an event e, a system variables valuation θ and a *ppDATE* variables valuation ν, predicate *nextState* holds whenever there exists an enabled transition on m going from q to q'. We formally write this as follows,

$$\begin{aligned} &\mathit{nextState}((m, q, \rho), e, \theta, \nu, q') \overset{df}{=} \\ &\quad \exists\, tr, c, a \cdot q \xrightarrow{tr|c\mapsto a}_{m} q' \text{ and} \\ &\qquad\qquad \mathit{activatedBy}(tr, e) \text{ and } \theta \cup \nu \models c \end{aligned}$$

$\square$

checkOnExit Given a local configuration (m, q, ρ), a system event $\sigma^{\downarrow}_{id}$, a system variables valuation θ, and a postcondition π', predicate *checkOnExit* holds if there exists a condition π such that the Hoare-triple $\{\pi\}\,\sigma\,\{\pi'\}$ is associated to state q, and π holds. We formally write this as follows,

$$\begin{aligned} &\mathit{checkOnExit}((m, q, \rho), \sigma^{\downarrow}_{id}, \theta, \pi') \overset{df}{=} \\ &\quad \exists\, \pi \cdot \{\pi\}\,\sigma\,\{\pi'\} \in \mathit{\Pi}_m(q) \text{ and } \theta \models \pi \end{aligned}$$

$\square$

enabled Given a local configuration l, an event e, a system variables valuation θ, and a *ppDATE* variables valuation ν, predicate *enabled* holds if either l has an enabled transition or it has a Hoare triple associated to q which has to be memorised. Formally,

$$\begin{aligned} &\mathit{enabled}(l, e, \theta, \nu) \overset{df}{=} \\ &\quad \exists\, q' \cdot \mathit{nextState}(l, e, \theta, \nu, q') \\ &\quad \text{or} \\ &\quad \exists\, \pi' \cdot \mathit{checkOnExit}(l, e, \theta, \pi') \end{aligned}$$

$\square$

toBeExecuted Given a local configuration (m, q, ρ), an event e, a system variables valuation θ, a *ppDATE* variables valuation ν, and an action a, predicate *toBeExecuted* holds if there exists an enabled transition such that a is its action. Formally,

$$\begin{aligned}
&\mathit{toBeExecuted}((m, q, \rho), e, \theta, \nu, a) \overset{df}{=} \\
&\quad \exists\ tr, c, q' \cdot \mathit{activatedBy}(tr, e) \text{ and} \\
&\qquad q \xrightarrow{tr|c\mapsto a}_m q' \text{ and } \theta \cup \nu \models c
\end{aligned}$$

□

Small Steps for Local Configurations The first step to define SOS rules describing the behaviour of a *ppDATE* network is to introduce rules showing how a local configuration performs a small step.

Given an event e, a system variables valuation θ, and a *ppDATE* variables valuation ν, a *small local configuration step* (or simply *small step local*), written $\xhookrightarrow{(e,\theta,\nu)}$, takes a local configuration (m, q, ρ) to some other local configuration (m, q', ρ'). This step relation is defined by the rules shown in Fig. 7. If e is an entry event of the form $\sigma^{\downarrow}_{id}$, there are three different possibilities: (i) there is an enabled transition in m going from state q to state q', and there is a Hoare triple $\{\pi\}\,\sigma\,\{\pi'\}$ associated to q such that π holds (entry_1); (ii) there is an enabled transition in m going from state q to q', but no Hoare triple $\{\pi\}\,\sigma\,\{\pi'\}$ associated to q such that π holds (entry_2); or (iii) there are no enabled transitions in m, but there is a Hoare triple $\{\pi\}\,\sigma\,\{\pi'\}$ associated to q such that π holds (entry_3).

In case of (entry_1), the next state reached by the enabled transition is q', and ρ gets extended by the tuple $(\sigma^{\uparrow}_{id}, \pi', \theta)$, in order to track the information about the postcondition which has to be checked upon the exit of method σ. Entry event identifiers are assumed to be unique in traces, and thereby, $\sigma^{\uparrow}_{id}$ is unique in ρ. In case of (entry_2) and (entry_3), only one of these two effects takes place. Then, apart from entry events, whenever e is either an exit event, i.e., it has the form $\sigma^{\uparrow}_{id}$, or an action event, by the rules *exit* and *act*, respectively, $\xhookrightarrow{(e,\theta,\nu)}$ results in the local configuration (m, q', ρ), where q' is the next state reached by the enabled transition.

Small Steps for Extended Global Configurations Given an extended global configuration $EC = (L, \nu, E, \theta)$, the relation *small step for extended global configurations* (or simply *small step global*), written as $\rightarrowtail$, takes EC to some extended global configuration (L', ν', E', θ) by following rule *iter*, which is depicted in Fig. 8. On this rule's premises, first of all, we define the set L_{en} of all the local configurations $(m, q, \rho) \in L$ such that m has an enabled transition whose triggers are activated by the events in E. Then, L_{en} is used to define both the set L_{nch} of local configurations in L that will *n*ot *ch*ange, and the set L_{ch} of the local configurations obtained

$$entry_1 \frac{\begin{array}{c} \mathit{checkOnExit}((m,q,\rho),\sigma^{\downarrow}_{id},\theta,\pi') \\ \mathit{nextState}((m,q,\rho),\sigma^{\downarrow}_{id},\theta,\nu,q') \end{array}}{(m,q,\rho) \xhookrightarrow{(\sigma^{\downarrow}_{id},\theta,\nu)} (m,q',\rho \cup \{(\sigma^{\uparrow}_{id},\pi',\theta)\})}$$

$$entry_2 \frac{\begin{array}{c} \nexists\ \pi' \cdot \mathit{checkOnExit}((m,q,\rho),\sigma^{\downarrow}_{id},\theta,\pi') \\ \mathit{nextState}((m,q,\rho),\sigma^{\downarrow}_{id},\theta,\nu,q') \end{array}}{(m,q,\rho) \xhookrightarrow{(\sigma^{\downarrow}_{id},\theta,\nu)} (m,q',\rho)}$$

$$entry_3 \frac{\begin{array}{c} \mathit{checkOnExit}((m,q,\rho),\sigma^{\downarrow}_{id},\theta,\pi') \\ \nexists\ q' \cdot \mathit{nextState}((m,q,\rho),\sigma^{\downarrow}_{id},\theta,\nu,q') \end{array}}{(m,q,\rho) \xhookrightarrow{(\sigma^{\downarrow}_{id},\theta,\nu)} (m,q,\rho \cup \{(\sigma^{\uparrow}_{id},\pi',\theta)\})}$$

$$exit \frac{\mathit{nextState}((m,q,\rho),\sigma^{\uparrow}_{id},\theta,\nu,q')}{(m,q,\rho) \xhookrightarrow{(\sigma^{\uparrow}_{id},\theta,\nu)} (m,q',\rho)}$$

$$act \frac{\begin{array}{c} e \in \mathsf{actevent} \\ \mathit{nextState}((m,q,\rho),e,\theta,\nu,q') \end{array}}{(m,q,\rho) \xhookrightarrow{(e,\theta,\nu)} (m,q',\rho)}$$

Fig. 7. Small Step Rules for Local Configurations

$$iter \frac{\begin{array}{c} L_{en} = \{l \mid l \in L, \mathit{enabled}(l,e,\theta,\nu), e \in E\} \\ L_{nch} = L \backslash L_{en} \\ L_{ch} = \{l' \mid l \in L_{en}, l \xhookrightarrow{(e,\theta,\nu)} l', e \in E\} \\ \mathit{Acts} = \{a \mid l \in L_{en}, \mathit{toBeExecuted}(l,e,\theta,\nu,a), e \in E\} \\ \mathit{mergeParalActs}_{\nu}(\{[\![a]\!]_{\theta,\nu} | a \in \mathit{Acts}\}) = (\nu', E', \mathit{New}') \\ L_{new} = \{(m, q_{0m}, \emptyset) \mid m \in \mathit{New}'\} \\ L' = L_{ch} \cup L_{nch} \cup L_{new} \end{array}}{(L,\nu,E,\theta) \rightarrowtail (L',\nu',E',\theta)}$$

Fig. 8. Small Step Rule for Extended Global Configurations

after performing a small step on the local configurations in L_{en}. We will use these two sets later to define L'. Next, we define the set *Acts* of all the actions which label the 'firing' transitions, and merge the meaning of those actions, which results in the valuation ν' and events E' of the new extended global configuration. We also initialise local configurations L_{new} for the newly created *ppDATE*s from *New'*. Finally, L' is the union of L_{ch}, L_{nch} and L_{new}.

$$\textit{shift}\,\frac{(L, \nu, \{e\}, \theta) \rightarrowtail^{*} (L', \nu', \emptyset, \theta)}{(L, \nu) \xRightarrow{(e,\theta)} (L', \nu')}$$

Fig. 9. Big Step Rules for Global Configurations

Note that if *mergeParalActs* is undefined, due to conflicts in parallel variable assignments (see Def. 20), then no global small step is defined, i.e., the execution aborts.

Big Steps for Global Configurations Given a *ppDATE* network $pn = (M, V, \nu_0, T_{ppd})$, a global configuration (L, ν) such that for all $(m, q, \rho) \in L$, $m \in M$ and $q \in Q_m$, and ν a valuation of the *ppDATE* variables V, a system event e and the system variables valuation θ, the relation *big step rules for global configurations* (or simply *big step global*), written $\xRightarrow{(e,\theta)}$, shifts (L, ν) to some global configuration (L', ν'), written $(L, \nu) \xRightarrow{(e,\theta)} (L', \nu')$, by rule *shift* given in Fig. 9. Note that, on this level, e and θ are external to the global configuration of the *ppDATE* network. Indeed, e and θ come from the system, and act as input to each step of the global configuration.

This rule means that whenever e occurs while the current system variables valuation is θ, (L, ν) shifts to (L', ν') if the transitive closure of relation *small step global* ($\rightarrowtail$, Fig. 8) takes the extended global configuration $(L, \nu, \{e\}, \theta)$ to the extended global configuration $(L', \nu', \emptyset, \theta)$. We need the transitive closure because the execution of actions may generate action events which also have to be consumed, meaning that we iterate using *small step global* until the set obtained by applying rule *iter* is the empty set. After having reached $(L', \nu', \emptyset, \theta)$, the small steps are saturated, because any configuration $(\llcorner\!\lrcorner, \llcorner\!\lrcorner, \emptyset, \llcorner\!\lrcorner)$ is a fixed-point of $\rightarrowtail$.

Lemma 1. *For each set of local configurations L, ppDATE variable valuation ν, and system variables valuation θ, the extended global configuration $(L, \nu, \emptyset, \theta)$ is a fixed-point of the relation small step global, i.e.,*

$$(L, \nu, \emptyset, \theta) \rightarrowtail (L, \nu, \emptyset, \theta)$$

Proof. In rule *iter* (Fig. 8), if $E = \emptyset$, then $L_{en} = L_{ch} = \mathit{Acts} = \emptyset$, and $L_{nch} = L$. From the note below Def. 20, we deduce that $(\nu', E', \mathit{New}') = (\nu, \emptyset, \emptyset)$, such that $L_{new} = \emptyset$, and $L' = L_{nch} = L$. Therefore, $(L', \nu', E', \theta) = (L, \nu, \emptyset, \theta)$.

□

We can now define the semantics of *ppDATE*s by identifying how a system trace changes the global configuration associated to a network of *ppDATE*s.

Definition 21. *We define how a system trace $w \in (systemevent \times \Theta_{Sys})^*$ shifts a ppDATE from the global configuration (L, ν) to the global configuration (L', ν'), written $(L, \nu) \overset{w}{\Rrightarrow} (L', \nu')$, by induction over w:*

$$(L, \nu) \overset{\varepsilon}{\Rrightarrow} (L', \nu') \overset{df}{=} L = L' \text{ and } \nu = \nu';$$
$$(L, \nu) \xRightarrow{w:(e,\theta)} (L', \nu') \overset{df}{=}$$
$$\exists\, L'', \nu'' \cdot (L, \nu) \overset{w}{\Rrightarrow} (L'', \nu'') \text{ and } (L'', \nu'') \xRightarrow{(e,\theta)} (L', \nu');$$

For this definition we will overload the operator we previously introduced to represent the relation *big step global*, i.e., $\Rrightarrow$ since it is straightforward to distinguish between the two from the context.

6.5 Valid Traces and Violating Traces

Before defining *violating system traces*, we have to introduce the notion of *counter-example.*

Definition 22. *Given a network of ppDATEs $pn = (M, V, \nu_0, T_{ppd})$, a system trace $w \in (systemevent \times \Theta_{Sys})^*$ is called a* counter-example *if $C_{init}(pn) \overset{w}{\Rrightarrow} (L, \nu)$, and (i) $\exists\ m, q, \rho \cdot (m, q, \rho) \in L \cdot q \in B_m$; or (ii) $w = w_1 +\!\!+ \langle(\sigma^{\uparrow}_{id}, \theta')\rangle$, $C_{init}(pn) \xRightarrow{w_1} (L', \nu')$ and $\exists\ m, q, \rho, \pi', \theta \cdot ((m, q, \rho) \in L'$ and $(\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho) \cdot \theta, \theta' \not\models \pi'$.* □

Note that (i) and (ii) are not exclusive, so a counter-example may have both properties at once.

Example 8. Recall the *ppDATE* m shown in Fig. 2. If m is in state q and event $\texttt{cleanF}^{\downarrow}_{1}$ occurs, the postcondition of $\{\texttt{true}\}\ \texttt{cleanF()}\ \{\texttt{cups == 0}\}$ is violated when method *cleanF* terminates. Thus, both $w = \langle(\texttt{brew}^{\downarrow}_{1}, \theta), (\texttt{brew}^{\downarrow}_{2}, \theta)\rangle$ and $w' = \langle(\texttt{brew}^{\downarrow}_{1}, \theta), (\texttt{brew}^{\uparrow}_{1}, \theta), (\texttt{cleanF}^{\downarrow}_{1}, \theta), (\texttt{cleanF}^{\uparrow}_{1}, \theta), (\texttt{brew}^{\downarrow}_{2}, \theta)\rangle$ are counter-examples. □

Definition 23. *The set of* violating system traces *of a ppDATE network pn, written $\mathcal{VT}(pn)$, is defined to be system traces which have a counter-example of pn as a prefix.* □

Definition 24. *The set of* valid system traces *of a ppDATE network pn, written $\mathcal{VAT}(pn)$, is defined to be the system traces which are not violating.* □

Example 9. The following system traces, for the coffee machine system of Fig. 2, are all valid:

$w = \langle(\texttt{brew}^{\downarrow}_{1}, \theta), (\texttt{brew}^{\uparrow}_{1}, \theta), (\texttt{brew}^{\downarrow}_{2}, \theta), (\texttt{brew}^{\uparrow}_{2}, \theta)\rangle$
$w' = \langle(\texttt{brew}^{\downarrow}_{5}, \theta), (\texttt{brew}^{\uparrow}_{5}, \theta), (\texttt{cleanF}^{\downarrow}_{2}, \theta), (\texttt{cleanF}^{\uparrow}_{2}, \theta)\rangle$
$w'' = \langle(\texttt{cleanF}^{\downarrow}_{4}, \theta), (\texttt{cleanF}^{\uparrow}_{4}, \theta), (\texttt{brew}^{\downarrow}_{2}, \theta), (\texttt{brew}^{\uparrow}_{2}, \theta)\rangle$ □

7 From *ppDATE* to *DATE*

In our framework (see Sec. 4), KeY first tries to prove all Hoare-triples of a *ppDATE* m, and then the partial proofs are used to get an optimised *ppDATE* m'. To make the property m' runtime-checkable, we further translate away the (remaining/optimised) Hoare triples, to arrive at a set of parallel (pure) *DATE*s that can be processed by Larva.

In this section, we formally define *DATE*s, we present the algorithm used by StaRVOOrS to translate *ppDATE*s into *DATE*s, finally, after introducing the semantics of *DATE*s, we prove soundness of the translation.

7.1 *DATE*

DATE [19] is a formalism similar to *ppDATE*, except that the automata do not include Hoare triples in the states. *DATE*s also include support for timers, which are not in *ppDATE*s. However, since the work we present here does not use timers, we leave them out from the formalisation.Formally:[9]

Definition 25. *A DATE is a ppDATE of the form* $(Q, t, B, q_0, \Pi_\emptyset)$*, where relation* $\Pi_\emptyset$ *represents that there are no Hoare triples assigned to any of the states in* Q*, i.e.,* $\Pi_\emptyset(q) = \emptyset$*,* $\forall q \in Q$*.*

□

Note that since a *DATE* is effectively a *ppDATE*, the semantics for *DATE*s are already covered by the semantics of *ppDATE*s. We will also refer to a (deterministic) network of *ppDATE*s where each *ppDATE* in the network is a *DATE*, as a network of *DATE*s and similarly *DATE* templates.

7.2 Translation from *ppDATE*s to *DATE*s

Here we present how to translate a *ppDATE* (network) into a *DATE* (network). However, first, let us intuitively analyse how the *ppDATE* depicted in Fig. 2, which we will refer to as m, can be translated into a *DATE* m'.

For simplicity, we assign the following names to the different Hoare triples in the states of m.

– h_1: {`cups` < `limit`} `brew()` {`cups == \old(cups)+1`}
– h_2: {`true`} `cleanF()` {`cups == 0`}

[9] Note that the definition of *DATE* given here is different from the one given in [19] as $\Pi_\emptyset$ was not defined in the original formulation. It is easy to see that the formulations are equivalent (modulo the differences mentioned above).

$exit\_cond\_checker = \lambda\ S, A : \Sigma, cond.$

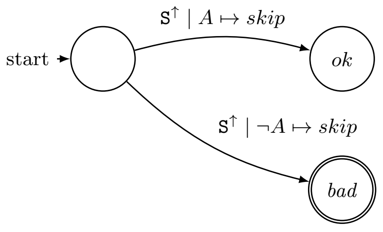

start
$\mathtt{S}^{\uparrow} \mid A \mapsto skip$
$\mathtt{S}^{\uparrow} \mid \neg A \mapsto skip$
ok
bad

Fig. 10. *DATE* template for verifying postconditions of Hoare triples.

- h_3: $\{$`cups` $<$ `limit`$\}$ `brew()` $\{$`cups == \old(cups)`$\}$
- h_4: $\{$`true`$\}$ `cleanF()` $\{$`cups == \old(cups)`$\}$

Then, we begin the translation by generating the *DATE* template illustrated in Fig. 10, which will be used to create *DATE*s in charge of controlling the postconditions of the previous Hoare triples.

Next, we start dealing with the translation of the transitions of m. m' will have exactly the same set of states as m, and it will have similar transitions to the ones of m. The only difference is that the transitions in m' will also have to address the verification of the Hoare triples. For instance, while being in state q, if the method `brew()` is executed and the precondition of h_1 holds, then its postcondition will have to be verified whenever method `brew()` finishes its execution.

Therefore, for every transition of the form $q \xrightarrow{\sigma^{\downarrow} | c \mapsto a}_{m} q'$, such that a Hoare triple $\{\pi\}\,\sigma\,\{\pi'\}$ is in q, m' will include a modified version of this transition in such a way that whenever this transition is fired, if π holds, then the execution of its action will have to create an instance of template *exit_cond_checker*. Thus, transitions t_1, t_3 and t_4 (recall Fig. 2) are modified as follows:

- $t_1' : q \xrightarrow{\mathtt{brew}^{\downarrow} | \mathtt{cups} < \mathtt{limit} \mapsto \mathsf{skip}\ ;\ a_1}_{m'} q'$
- $t_3' : q' \xrightarrow{\mathtt{brew}^{\downarrow} | \mathsf{true} \mapsto \mathsf{skip}\ ;\ a_2}_{m'} bad$
- $t_4' : q' \xrightarrow{\mathtt{cleanF}^{\downarrow} | \mathsf{true} \mapsto \mathsf{skip}\ ;\ a_3}_{m'} bad$

where,

- a_1 : if (`cups` $<$ `limit`)
  then *create(exit_cond_checker,*`brew`*,*`part_eval(cups==\old(cups)+1)`*)*
- a_2 : if (`cups` $<$ `limit`)
  then *create(exit_cond_checker,*`brew`*,*`part_eval(cups==\old(cups))`*)*
- a_3 : if (`true`)
  then *create(exit_cond_checker,*`cleanF`*,*`part_eval(cups == \old(cups))`*)*

In the previous transitions we have used as the conditions of the if-expressions in actions a_1, a_2 and a_3, the preconditions of the different Hoare triples to be verified in each case. Moreover, function `part_eval` partially evaluates its argument, replacing the expressions `\old`(e) operator the current value of e. If a postcondition does not include such operator, then `part_eval` is the identity. Note that even though the if-expression in transition t'_4 may seem unnecessary, we include it anyway in order to exactly reflect how the translation algorithm works.

In addition, if at a certain state, a Hoare triple has to be verified, but in that state there are no outgoing transitions with an event related to the method in the Hoare triple, then a new transition is added to m' in order to be able to control such Hoare triple. For instance, in state q the following *self*-transition has to be added in order to verify h_2 and h_3.

– $t'_5 \colon q \xrightarrow{\texttt{cleanF}^{\downarrow}|\texttt{true}\mapsto a_4}_{m'} q$

where,

– a_4: *create(exit_cond_checker,*`cleanF`*,*`part_eval(cups == 0)`*)*

Again, we use the preconditions of the Hoare triples as conditions of the previous action.

Given a transition $q \xrightarrow{tr|c\mapsto a}_{m} q'$ such that (i) tr fires upon exiting a method, or (ii) tr fires upon entering a method but there is no Hoare triple associated to this method in q, these transitions remain untouched, i.e., it is translated as $q \xrightarrow{tr|c\mapsto a}_{m'} q'$. For instance, transition t_2 is translated as follows.

– $t'_2 : q' \xrightarrow{\texttt{brew}^{\uparrow}|\texttt{true}\mapsto skip}_{m'} q$

Fig 11 illustrates the *DATE* obtained when translating m following the previous steps (i.e., m'). Note that whole translation would consist on the previous *DATE* and the generated template *exit_cond_checker*.

Translation Algorithm For clarity of presentation we give two algorithms, one for the case when no Hoare triples clashes arise, and one for the full case. Intuitively, we call it a clash if the behaviour of a method σ, in a certain *ppDATE* state q, is defined by both, a Hoare triple in q, and an outgoing transition from q. Formally, we define a clashing Hoare triple as follows.

Definition 26. *Given a ppDATE network* $pn = (M, V, \nu_0, T_{ppd})$ *such that every ppDATE* $m \in M$ *is defined as the tuple* $(Q_m, t_m, B_m, q_{0m}, \Pi_m)$*, a Hoare triple* $\{\pi\}\,\sigma\,\{\pi'\} \in \Pi_m(q)$*, for some* $q \in Q_m$*, is called* clashing *if an outgoing transition from* q *is guarded by trigger* $\sigma^{\downarrow}$ *(i.e.,* $\exists\ c, a, q' \cdot q \xrightarrow{\sigma^{\downarrow}|c\mapsto a}_{m} q'$*). A* clash-free *ppDATE is a ppDATE with no clashing Hoare triples.* □

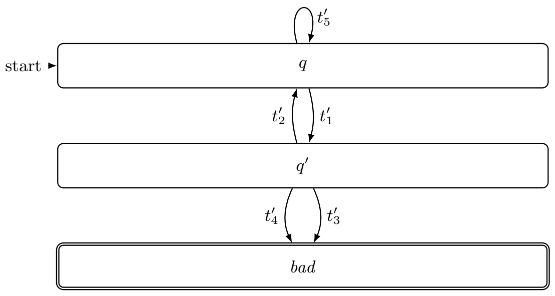


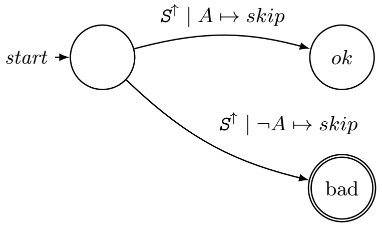
Fig. 11. Translation to *DATE* of the *ppDATE* illustrated in Fig. 2.

We now present the algorithm to translate a clash-free *ppDATE* network into a *DATE* network. The translation works by replacing each Hoare triple $\{\pi\}\,\sigma\,\{\pi'\}$ in a state q of a *ppDATE* by a new reflexive transition (from q to q) triggered by an entry into function σ such that the precondition π holds, and a parallel *DATE* is created, checking the postcondition.

We assume a function `part_eval` $\in$ *postcond* $\mapsto$ *cond*, which removes `\old` constructs in postconditions. The function performs *partial evaluation* — replacing each `\old`(e) with the current value of e. Our algorithm syntactically places the `part_eval` function in an action that will be executed when the according method is entered, i.e., partial evaluation does not happen during the translation algorithm, but at runtime, when the method is entered.

Algorithm 1. *Given a* clash-free *ppDATE network* $pn = (M, V, \nu_0, T_{ppd})$, *such that every ppDATE* $m \in M$ *is defined as the tuple* $(Q_m, t_m, B_m, q_{0m}, \Pi_m)$*, we can construct a DATE network equivalent to pn in the following manner:*

1. *With each Hoare triple* $\{\pi\}\,\sigma\,\{\pi'\}$ *in a ppDATE state, replace in* π' *each instance of the* `\result` *by the variable* `ret`*. This variable will represent the value returned by the method associated to the Hoare triple/*
2. *Generate the following DATE template:*
   exit_cond_checker $= \lambda\ S, A : \Sigma, cond.$


$S^{\uparrow} \mid A \mapsto skip$
start
ok
$S^{\uparrow} \mid \neg A \mapsto skip$
bad

This template will be used to create DATEs handling the verification of the postcondition of the method.

3. *Transform M, the set of ppDATEs of pn, into the set of DATEs $M' = \{m' \mid m' = (Q_m, t'_m, B_m, q_{0m}, \Pi_\emptyset), m \in M\}$ such that t'_m follows the rules below:*

   3a. each Hoare triple $\{\pi\}\,\sigma\,\{\pi'\}$ in $\Pi_m(q)$ is replaced by $q \xrightarrow{\sigma^{\downarrow}|\pi\mapsto a}_{m'} q$, where $a = \mathsf{create}(exit\_cond\_checker, \sigma, \mathrm{part\_eval}(\pi'))$;

   3b. each transition $q \xrightarrow{tr|c\mapsto a}_{m} q'$ remains unchanged, i.e. $q \xrightarrow{tr|c\mapsto a}_{m'} q'$

4. *Translate T_{ppd} (the set of ppDATE templates in pn) into a set of DATE templates T_d by repeatedly applying step 3a. and 3b. to the body of templates.*
5. *Extend the set T_d by including the template generated in step 2. Let us call this extension T'_d.*
6. *Finally, the resulting DATE network is defined to be (M', V, ν_0, T'_d).*

This translation works well except that it would introduce non-determinism when the *ppDATE* includes clashes. To extend the translation to work in the presence of clashes, we transform Hoare triples clashing with a transition into a family of disjoint transitions, each of which performs the transition but also checks whether the postcondition checker should be created.

Algorithm 2. *Given a (possibly clashing) ppDATE network pn, we construct a network of DATEs equivalent to pn by using Algorithm 1 except that we replace steps 3.a and 3.b, by the following:*

$3a_1$. Each non-clashing Hoare triple: $\{\pi\}\,\sigma\,\{\pi'\}$ in $\Pi_m(q)$ is turned into a transition $q \xrightarrow{\sigma^{\downarrow}|\pi\mapsto \mathsf{create}(exit\_cond\_checker,\sigma,\mathrm{part\_eval}(\pi'))}_{m'} q$

$3a_2$. For each clashing Hoare triple: $\{\pi\}\,\sigma\,\{\pi'\} \in \Pi(q)$, clashing with n outgoing transitions, $q \xrightarrow{\sigma^{\downarrow}|c_k\mapsto a_k} q^k$ $(0 \leq k < n)$:

- *Replace $q \xrightarrow{\sigma^{\downarrow}|c_k\mapsto a_k}_{m} q^k$ with: $q \xrightarrow{\sigma^{\downarrow}|c_k\mapsto(a_k;\, \mathit{if}\,\pi\ \mathbf{then}\, a)}_{m'} q^k$;*
- *Add the following transition: $q \xrightarrow{\sigma^{\downarrow}|(!c_0\&\&\ldots\&\&!c_n\&\&\pi)\mapsto a}_{m'} q$,*

where, in both cases, $a = \mathsf{create}(exit\_cond\_checker, \sigma, \mathrm{part\_eval}(\pi'))$

3b. each transition $q \xrightarrow{tr|c\mapsto a}_{m} q'$ such that either $\Pi_m(q) = \emptyset$, $\Pi_m(q) \neq \emptyset$ but there is no Hoare triple associated to trigger tr, or trigger tr is activated by an exit event, remains unchanged, i.e. $q \xrightarrow{tr|c\mapsto a}_{m'} q'$.

7.3 Proof of Soundness of the translation algorithm

In this section we will show that the translation algorithms introduced in the previous section are sound.

7.4 Coupling Invariant Lemmas

Here, we formally introduce two lemmas which together form the coupling invariant that is used to prove soundness. The proofs of these lemmas can be found in Appendix A.

Lemma 2 states that given a trace, both a *ppDATE* network *pn* and its translation to *DATE* will change their initial global configuration to global configurations (L, ν) and $(\tilde{L}, \nu')$, respectively, such that $\nu = \nu'$, and that for every $(m, q, \rho) \in L$ where m is in *pn*, there is a local configuration $(m', q', \emptyset) \in \tilde{L}$ such that m' is the translation of m and both m and m' are in the same state, and vice versa.

In this lemma we represent the translation of a single *ppDATE* to *DATE* with the function $\kappa \in ppDATE \mapsto DATE$.

Lemma 2. *Given a network of ppDATEs* $pn = (M, V, \nu_0, T_{ppd})$*, its translation* $ppd2DATE(pn) = (M', V, \nu_0, T'_d)$*, a trace* $w \in (systemevent \times \Theta_{Sys})^*$*, and the global configurations* (L, ν) *and* $(\tilde{L}, \nu')$*,*

$$
\begin{array}{l}
C_{init}(pn) \overset{w}{\Rightarrow}_M (L, \nu) \text{ and } C_{init}(ppd2DATE(pn)) \overset{w}{\Rightarrow}_{M'} (\tilde{L}, \nu') \\
\text{implies} \\
\quad \nu = \nu' \\
\quad \text{and} \\
\quad \forall\ m, q, \rho \cdot (m, q, \rho) \in L, m \in M \cdot \\
\qquad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}\ \cdot \kappa(m) = m' \textit{ and } q = q' \\
\quad \text{and} \\
\quad \forall\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \in M' \cdot \\
\qquad \exists\ m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L \cdot q = q' \\
\quad \text{and} \\
\quad \forall\ m, q, \rho \cdot (m, q, \rho) \in L, m \notin M \cdot \\
\qquad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot q = q' \\
\quad \text{and} \\
\quad \forall\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot \\
\qquad \exists\ m, q, \rho \cdot (m, q, \rho), m \notin M \in L \cdot q = q'
\end{array}
$$

Lemma 3 states that given a trace, if this trace shifts a *ppDATE* network *pn* and its *DATE* translation from their respective initial global configuration to some global configurations (L, ν) and $(\tilde{L}, \nu')$, respectively, then for each entry $(\sigma^{\uparrow}_{id}, \pi', \theta)$ in a ρ component of a local configuration in L there is one local configuration in $\tilde{L}$ whose *DATE* component is an instance of the template *exit_cond_checker* in charge of controlling π', and vice versa.

Lemma 3. *Given a network of ppDATEs* $pn = (M, V, \nu_0, T_{ppd})$*, its translation* $ppd2DATE(pn) = (M', V, \nu_0, T'_d)$*, a trace* $w \in (systemevent \times \Theta_{Sys})^*$*, and the global configurations* (L, ν) *and* $(\tilde{L}, \nu')$*,*

$$C_{init}(pn) \overset{w}{\Rightarrow}_M (L, \nu) \text{ and } C_{init}(ppd2DATE(pn)) \overset{w}{\Rightarrow}_{M'} (\tilde{L}, \nu') \text{ implies } \psi(L, \tilde{L})$$

where,

$$
\begin{aligned}
\psi(L,\tilde{L}) = {} & \forall\ m,q,\rho \cdot (m,q,\rho) \in L\ \cdot \\
& \quad \forall\ \sigma^{\uparrow}_{id},\pi',\theta \cdot (\sigma^{\uparrow}_{id},\pi',\theta) \in \rho\ \cdot \\
& \quad\quad \exists\ m',q' \cdot (m',q',\emptyset) \in \tilde{L} \cdot m' = \mathit{inst}\,(\mathit{exit\_cond\_checker},\sigma,\pi') \\
& \mathit{and} \\
& \quad \forall\ m',q' \cdot (m',q',\emptyset) \in \tilde{L}, m' \notin M' \cdot \\
& \quad\quad \exists\ \sigma^{\uparrow}_{id},\pi' \cdot m' = \mathit{inst}\,(\mathit{exit\_cond\_checker},\sigma,\pi') \\
& \quad\quad\quad \text{implies } \exists\ m,q,\rho,\theta \cdot (m,q,\rho) \in L \cdot (\sigma^{\uparrow}_{id},\pi',\theta) \in \rho
\end{aligned}
$$

□

Proof of Soundness We can now prove the soundness of the translation algorithm. Below we provide the formalisation of this property and an intuitive explanation for it. However, a rigorous proof of this theorem can be found in Appendix B.

Theorem 1. *Given a ppDATE network $pn = (M, V, \nu_0, T_{ppd})$, and its translation ppd2DATE(pn) $= (M', V, \nu_0, T'_d)$,*

$$\mathcal{VT}(pn) = \mathcal{VT}(ppd2DATE(pn))$$

Proof. To prove the soundness of the translation algorithm we will show that both a *ppDATE* network *pn* and its translation to a *DATE* network have the same set of violating traces. Intuitively, we will prove that given a trace w which is violating for *pn*, i.e., $w \in \mathcal{VT}(pn)$, is also violating for *pn*'s translation, i.e., $w \in \mathcal{VT}(ppd2DATE(pn))$, and vice versa.

In the case when $w \in \mathcal{VT}(pn)$, by definition of counter-examples of *ppDATE*s, w has a prefix w' such that either (i) w' takes the initial global configuration $C_{init}(pn)$ to a global configuration (L', ν') such that the state component of L' is a bad state; (ii) given a method σ and a system variables valuation θ', w' can be written as $w_1 +\!\!+ (\sigma^{\uparrow}_{id}, \theta')$ such that w_1 takes $C_{init}(pn)$ to a global configuration (L', ν') where there exists a local configuration in L' whose ρ component contains a tuple $(\sigma^{\uparrow}_{id}, \pi', \theta)$, such that π' fails to be satisfied in the 'moment' event $\sigma^{\uparrow}_{id}$ appears.

In the case of (i), we use the fact that (by Lemma 2), if w' takes the translation from the initial global configuration $C_{init}(ppd2DATE(pn))$ to a global configuration $(\tilde{L}, \nu)$, for every local configuration in L', there is a local configuration in $\tilde{L}$ such that its state component is the same. Thus, there is a local configuration in $\tilde{L}$ whose state component is a bad state, which means that w' is a counter-example of the translation as well.

In the case of (ii), due to the fact that a Hoare triple $\{\pi\}\,\sigma\,\{\pi'\}$ has to be verified, we know that some local configuration will have a ρ component such that $(\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho$. We can now use the fact that by Lemma 3, tuple is handled by a *DATE* in the translation (which verifies the postcondition).

Thus, there exists a *DATE* controlling π' which fails moving to a bad state, i.e., w' is a counter-example of the translation as well.

In order to prove the opposite direction, we proceed to assume that $w \in \mathcal{VT}(\mathit{ppd2DATE}(pn))$. Again, since this is a counter-example and this is a *DATE* (and thus cannot fail due to a violated postcondition), it can be only the case that w has a prefix w' such that this prefix takes the initial global configuration $C_{init}(\mathit{ppd2DATE}(pn))$ to a global configuration $(\tilde{L}, \nu)$ such that there is a local configuration in $\tilde{L}$ whose state component is a bad state. Then, assuming that w' takes pn from the initial global configuration $C_{init}(pn)$ to a global configuration (L', ν'), we proceed to do a case analyses depending whether the bad state belongs to a *DATE* which was controlling the postcondition of a Hoare triple or not. In the affirmative case, we will use this fact to show that, given certain method σ and a system variables valuation θ', w' can be selected to be a prefix which can be written as $w_1 +\!\!+ (\sigma^{\uparrow}_{id}, \theta')$ such that w_1 takes $C_{init}(pn)$ to a global configuration (L', ν') where the verification of the postcondition fails whenever event $\sigma^{\uparrow}_{id}$ occurs. Therefore, w' is a counter-example of pn. Finally, (by Lemma 2), there is a local configuration in L' such that its state component is the same as the bad state in $\tilde{L}$. Therefore, w' is a counter-example of pn. □

8 The StaRVOOrS Tool Implementation

In this section we present how the (fully automatic) verification tool StaRVOOrS [17] implements the framework presented in Sec. 4. To illustrate this, we use a running example of a *bank system* in which users log in to perform transactions[10]. The set of logged-in users is implemented as a `Hashtable` object, whose class represents an open addressing hashtable with linear probing as collision resolution. Method `add`, which is used to add objects into the hashtable, first attempts to put the corresponding object at the position of its computed hash code. However, if that index is occupied, then `add` searches for the nearest following index which is free. Fig. 12 depicts a code snippet for this method. Within the hashtable object, users are stored into an array `arr`. This means that the set of logged-in users has its capacity limited by the length of `arr`. In order to check in a straightforward manner whether the capacity of `arr` is reached or not, a field `size` keeps track of the amount of stored objects and a field `capacity` represents the (total) number of objects that can be added into the hash table. In addition, this system has to fulfil the properties described with the *ppDATE* template depicted in Fig. 13. This template specifies the following properties:

[10] Both the source code and the *ppDATE* specification for this example are available from [4].

```
public void add (Object o, int key) {
    if (size < capacity) {
       int i = hash_function(key);
       if (h[i] == null) {
          h[i] = o;
          size++;
          return;
       }
       else {
          while (h[i] != null) {
                if (i == capacity-1) i = 0;
                else {i++;}
          }
          h[i] = o;
          size++;
          return;
       }
    }
}
```

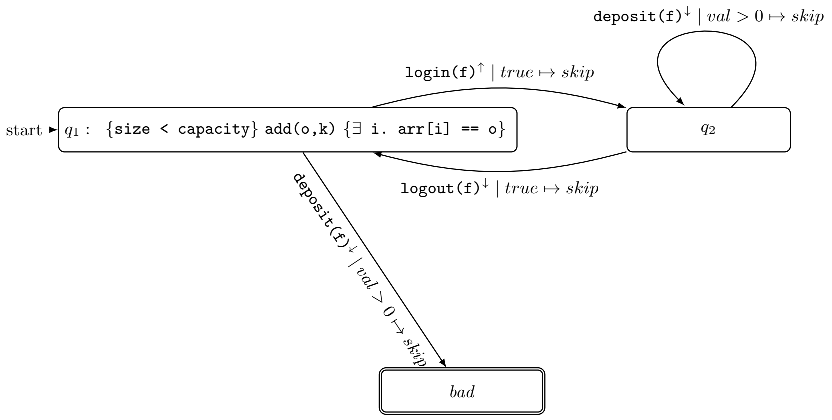
Fig. 12. Code snippet for method `add`.

(i) *A user has to be logged-in in order to perform a deposit, i.e. a deposit should happen between a login and a logout.*

(ii) *Provided there is space in the hashtable, executing method* `add` *with object* `o` *and key* `k` *should add the object to the table.*

Property (i) is verified with the transitions of the *ppDATE* template, whereas property (ii) is represented by the Hoare triple in state q_1. If `size < capacity`, then there is room in the hashtable for one more element, and if method `add` places the object `o` in the hashtable, there exists an index in the array `arr` such that `o` is placed in that index, i.e., $\exists$ `int i; i>= 0 && i < capacity; arr[i] == o`. Note that the given Hoare triple is only included in state q_1 since only a successful login leads to the execution of the method `add`, i.e., this Hoare triple is context dependent; and that `login(f)`$^{\downarrow}$ means that method `login` associated to the trigger is the one defined within object f. In addition, we assume that the specification of the system has a *ppDATE* with a single state q and single transition of the form $q \xrightarrow{\texttt{new(o)}^{\downarrow}|\texttt{true}\mapsto create(prop{-}deposit{-}temp,\texttt{o})} q$, such that the trigger `new(o)`$^{\downarrow}$ is activated by the declaration of an object o of the class `UserInterface`. Thus, this *ppDATE* creates an instance of the template in Fig. 13 every time an object of the class `UserInterface` is declared.

prop-deposit-temp = $\lambda\ f$: `UserInterface`.



deposit(f)↓ | val > 0 ↦ skip
login(f)↑ | true ↦ skip
start
q1 : {size < capacity} add(o,k) {∃ i. arr[i] == o}
q2
logout(f)↓ | true ↦ skip
deposit(f)↓ | val > 0 ↦ skip
bad

Fig. 13. *ppDATE* specification of properties for a bank system.

8.1 *ppDATE* Specification as an Input Script for StaRVOOrS

Before describing how StaRVOOrS works, we need to introduce how a *ppDATE* specification is written as an input script for this tool. Below, we show the input script for the *ppDATE* template illustrated in Fig. 13, and the *ppDATE* which creates its instances. In addition, we give a brief description of each one of the sections this script. For a full description on how to write *ppDATE*s as an input script for our tool, one may refer to the StaRVOOrS *User Manual*[11].

```
IMPORTS { main.UserInterface ; main.Hashtable ; }

GLOBAL {
  PROPERTY prop-deposit {
      PINIT { (prop-deposit-temp, UserInterface) }
  }
}
TEMPLATES {
 TEMPLATE prop-deposit-temp (UserInterface uf) {
   TRIGGERS {
     login_exit(String un, int pwd)
        = {UserInterface f.login(un, pwd)exit()} where {uf = f}
     logout_entry()
        = {UserInterface f.logout()entry} where {uf = f}
     deposit_entry(int val)
```

[11] This document is available from [4], in the Downloads section.

```
        = {UserInterface f.deposit(val)entry} where {uf = f}
   }
   PROPERTY prop_deposit {
     STATES {
       ACCEPTING { q2 ; }
       BAD { bad ; }
       STARTING { q1 (add_ok) ; }
     }
     TRANSITIONS {
       q1 -> q2 [login_exit \ f.getUser() != null]
       q1 -> bad [deposit_entry]
       q2 -> q1 [logout_entry \ f.getUser() != null ]
       q2 -> q2 [deposit_entry \ f.getUser() != null]
     }
   }
  }
}
CINVARIANTS {
  HashTable {\typeof(h) == \type(Object[])}
  HashTable {arr.length == capacity}
  HashTable {arr != null}
  HashTable {size >= 0 && size <= capacity}
  HashTable {capacity >= 1}
}
HTRIPLES {
  HT add_ok {
    PRE {size < capacity}
    METHOD {Hashtable.add}
    POST  {(\exists int i; i>= 0 && i < capacity; arr[i] == o)}
    ASSIGNABLE {size, arr[*]}
  }
}
```

The section `IMPORTS` lists the Java packages which may be used in any of the other sections of the script, in this case `UserInterface` and `Hashtable`. The section `TEMPLATES` contains the description of the *ppDATE* templates (tagged by `TEMPLATE`). Here, the section `TRIGGERS` is used to declare the different triggers which may be used in the transitions of the *ppDATE*, i.e, `login_exit`, `logout_entry`, `deposit_entry`, and the section `PROPERTY` describes the different states, i.e., `q1`, `q2` and `bad`, and transitions of the *ppDATE*. Note that the syntax `q1 (add_ok)` associates the Hoare triple tagged as `add_ok` to state `q1`. This means that the Hoare triple `add_ok` has to be verified if the method associated to it, in this case method `add`, is executed whenever the *ppDATE* is in state `q1`. The section `GLOBAL` contains the description of the *ppDATE*. Here, *ppDATE*s are described in the same manner as in a `TEMPLATE` section. However, note that it is also possible, as it is the case in our example, to use the special section `PINIT` when describing the section `PROPERTY`. Section `PINIT` represents a *ppDATE* with single state, and a looping transition which is fired every time an object of the class listed within this section (`UserInterface` in our example) is declared, leading to the creation of an instance of the listed template for that object (*prop-deposit-temp* in our example). We have included this special case because it is quite common to have *ppDATE*s only focus on creating instances of a template upon declaration of a particular object. Regarding the section `CINVARIANTS`, class invariants are described by the syntax `class_name` {`invariant`}, meaning that `invariant` has to be fulfilled by all the methods in the class `class_name`. These invariants are only meant as a help for the deductive verification of the Hoare triples (see Sec. 8.2). If no invariants are needed, then this section can be omitted. Finally, the section `HTRIPLES` gives a list of named Hoare triples (tagged by `HT`). Here, `PRE` describes the precondition of the Hoare triple, `POST` describes the postcondition of the Hoare triple, `METHOD` indicates which one is the method associated to the Hoare triple, and `ASSIGNABLE` lists the (class) variables that might be modified when the method associated to the Hoare triple is executed. Note that the predicates in invariants, pre- and postconditions follows JML-like syntax and pragmatics. For instance, in the Hoare triple `add_ok` the second semicolon separates the range predicate (`i>=0 && i<capacity`) from the desired property over integers in that range, (`arr[i]==o`).

8.2 Running StaRVOOrS

StaRVOOrS is a fully automatic verification tool which takes the Java source code of the system under scrutiny and a file with the *ppDATE* specification for this system and produces (i) a runtime monitor, (ii) an instrumented version of the system given as input with event generation and additional code infrastructures required, (iii) a report summarising the results of the deductive verification of the Hoare triples, and (iv) a refined version (if any) of the provided *ppDATE* specification.

This tool implements the framework described in Sec. 4 with each stage of the framework, i.e., *Deductive Verification*, *Specification Refinement*, *Translation and Instrumentation*, and *Monitor Generation*, being performed automatically by the tool. Below, we describe the implementation of these stages through the use of the working example.

Deductive Verification The first step performed by StaRVOOrS is the deductive verification of the Hoare triples associated to the states of the *ppDATE* (template) using KeY. To accomplish this, StaRVOOrS extracts the Hoare triples specified in the *ppDATE* script, converts them into JML contracts, and then annotates these contracts in the Java sources, before the corresponding method declaration. For instance, the following JML contract associated to method `add` is extracted from the Hoare triple `add_ok`:

```
requires size < capacity;
ensures (\exists int i; i>= 0 && i < capacity ; arr[i] == o);
assignable size, arr[*];
```

Note that the `requires` clause describes the precondition of `add`, the `ensures` clause describes the postcondition of `add`, and the `assignable` clause lists the (class) variables that might be modified when `add` is executed.

Once all the JML contracts are in place, i.e., they are annotated in the code, StaRVOOrS uses KeY to verify them. First, KeY generates proof obligations in Java Dynamic Logic for each JML contract. Next, it attempts to prove the contracts automatically. Finally, it stores the results of all the verification attempts in a XML file. Here, note that even though it could be possible to allow for user interaction (using KeY's elaborate support for interactive theorem proving), we chose to use KeY in automatic mode, since StaRVOOrS targets users untrained in theorem proving. StaRVOOrS generates a report summarising the results produced by KeY in an easy to understand format.

Using our running example, when KeY tries to verify the previous JML contract, it will result in a partial proof. This analysis is shown in the following fragment of the generated XML file:

```
<executionPath
  pathCondition="arr[hash_function(key)] = null"
  verified="true"/>
<executionPath
  pathCondition="!arr[hash_function(key)] = null"
  verified="false"/>
```

This indicates that while KeY was symbolically executing method `add`, there was a branching in the condition `arr[hash_function(key)] = null`, leading to two possible execution paths (depending on its truth value). Recalling the code snippet in Fig. 12, this condition corresponds to the condition on the if-expression in line 4. Thus, the execution path for the condition `arr[hash_function(key)] = null` corresponds to the case where the array `arr` has a free slot at the hash code of `key`, whereas the execution path for the condition `!arr[hash_function(key)] = null` corresponds to the case where the program enters the while-loop in line 10, searching for the next free slot in `arr`. In addition, in the XML, the component `verified` represents whether KeY was able to prove the branch of the proof (`verified=true`), or not (`verified=false`). Therefore, from the previous fragment of the XML file we know that KeY was able to close the branch where the array `arr` has a free slot (`= null`) at the hash code of `key`, but it was not able to verify the other case (where the program enters a loop searching for the next free slot). The main reason why KeY was not able to prove the latter case is the lack of loop invariants to deal with the while-loop.

Specification Refinement The output pf KeY is then used to refine the Hoare triples in the specification based on what was (partially) proved. The Hoare triples associated to JML contracts which were fully verified by KeY are entirely removed from the specification, while the precondition of the Hoare triples associated to partially proved JML contracts are refined based on what KeY managed to prove. The new precondition is the conjunction of the original precondition with the disjunction of new preconditions corresponding to open proof goals, i.e., the path condition on each different execution paths. Note that StaRVOOrS generates a new *ppDATE* specification script based on such refinements, instead of modifying the provided *ppDATE* script.

In the example, the precondition of the Hoare triple `add_ok` will be refined with the condition for the one goal not closed by KeY, i.e., `!(arr[hash_function(key)] == null)`. The Hoare triple will thus be strengthened as follows:

```
HT add_ok {
  PRE {size < capacity && !(h[hash_function(key)] == null)}
  METHOD {Hashtable.add}
  POST {(\exists int i; i>=0 && i<capacity; arr[i]==o)}
  ASSIGNABLE {size, arr[*]}
}
```

Translation and Instrumentation Once the refined *ppDATE* specification is ready, StaRVOOrS translates it into (pure) *DATE* formalism using the algorithm from Sec.7.2. This enables the monitor generation by Larva (explained in the next stage). In addition, in order to properly address the refined *ppDATE*, our tool operationalise the conditions and instruments the code, as described below.

Pre/Postcondition Operationalisation In this step, the tool syntactically analyses the specification for expressions in pre- and postconditions of the Hoare triples which may have to be operationalised, i.e., transformed into algorithmic procedures. For instance, transforming either existential or universal quantifications into loops.

During the operationalisation process, the tool creates Java code containing the implementation of all necessary methods for runtime verification, including those generated to algorithmically check the pre/postconditions.

In our example, as the postcondition of the Hoare triple `add_ok` has an existential quantifier, it has to be operationalised, producing the following method:

```
public static boolean add_ok_post_opE_1(Hashtable hasht,
     Object o, int key) {
  boolean r = false;
4   for (int i = 0 ; i < hasht.capacity ; i++) {
5     if (hasht.arr[i] == o) { r = true ; break; }
6   }
7   return r;
8 }
```

The for-loop declaration in line 3 is created from the conditions in the range of the existential quantification, i.e., `i>=0 && i<capacity`, and the condition of the if-expression in line 4 is created from the condition in the body of the existential quantification, i.e., `arr[i]==o`. Thus, if any value on the range of the existential quantification fulfils its body, then this method returns `true`, i.e., exists a value that fulfils the existential quantification. Otherwise, it returns `false`, i.e., it does not exist a value fulfilling the existential quantification.

Code Instrumentation Next, StaRVOOrS instruments the Java source code of the system adding identifiers to each method associated to a Hoare triple in the refined *ppDATE* specification script, and additional code to get fresh identifiers. As mentioned in Sec. 4, these identifiers will be used to distinguish different executions of the same method. However, in order to avoid modifying all the calls to these methods in the entire system, we have opted to introduce this instrumentation in the form of auxiliary methods. For instance, in our working example the method `add` has to be instrumented, resulting in:

```
public void add (Object o, int key) {
   addAux(o,key,fid.getNewId());
}
public void addAux (Object o, int key, Integer id) {...}
```

The method `addAux` implementation corresponds to the body of method `add` in Fig. 12. This method represents the instrumentation of method `add` with the extra argument `Integer id`, which is used as identifier. In addition, method `add` now simply calls `addAux`, but generating a fresh identifier for the call using function `fid.getNewId`.

Moreover, the previously generated *DATE* specification is modified accordingly, to refer to the instrumented version of the methods. In our example, the *DATE* specification would be modified to refer to method `addAux` instead of method `add`.

Monitor Generation Finally, StaRVOOrS uses Larva to automatically generate a monitor from the *DATE* specification obtained in the previous stage. Larva takes this *DATE* and generates the monitoring system and aspects instrumenting the communication between the system and the monitor [20].

9 Case Study: SoftSlate Commerce

SoftSlate Commerce (or simply SoftSlate) [3] is an open-source Java shopping cart web application designed following a *Model-View-Controller* architecture. A user of SoftSlate sends a request to a server hosting the application via a web browser. Then, the server processes the received request and executes an action associated to it (*Controller layer*). Such action may require to interact with and/or modify the information in the database (*Model layer*), e.g., information about users, products, orders, etc. Finally, once the request is fully processed, the server sends back a response to the user. The information in this response will be reflected on a web page loaded on the browser (*View layer*). The administrator of the application interacts with it in a similar fashion.

SoftSlate offers a basic implementation of a shopping cart web application featuring outer space related pictures, whose server is set up by using *Apache Tomcat* [1]. This implementation is meant to be used by developers to start building their own web applications.

In this case study we analyse an extension of the SoftSlate basic implementation. This extension increases modularity of parts of the implementation, to better link it to the required properties. Basically, we have created a few helper methods in order to better observe the various steps performed by a user to checkout a purchase. In addition, we have modified a few methods to receive an entire object instead of some of its components, and to properly access the components.

As our main focus is to verify the source code offered by SoftSlate, in our extension we are not adding any new feature to the ones already provided in the basic implementation, i.e., the functionality of the basic implementation and our extension is the same.

Note that when this case study was developed, there was not an open source version of SoftSlate available online, meaning that we cannot distribute the sources we have used. However, in [4] one may find the files for the *ppDATE* specifications described below.

9.1 *ppDATE* specification

Here we introduce two *ppDATE*s specifications, one describing a property related to the log in and log out of users in the web application, and one describing a property related to the checkout of the purchases performed by the users of the application. These properties address basic functionalities which we consider that a web cart application should offer.

Note that even though we could have either described more properties or specified more control- and data-oriented behaviour in the properties we are depicting in this section, the *ppDATE*s introduced here are sufficient to highlight the benefits of using StaRVOOrS in a real application. In addition, for readability reasons, Hoare triples are not going to be included

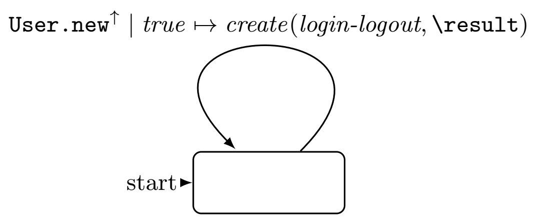


Fig. 14. *ppDATE* in charge of creating instances of the template *login-logout*.

login-logout $= \lambda\ u$: `User`.

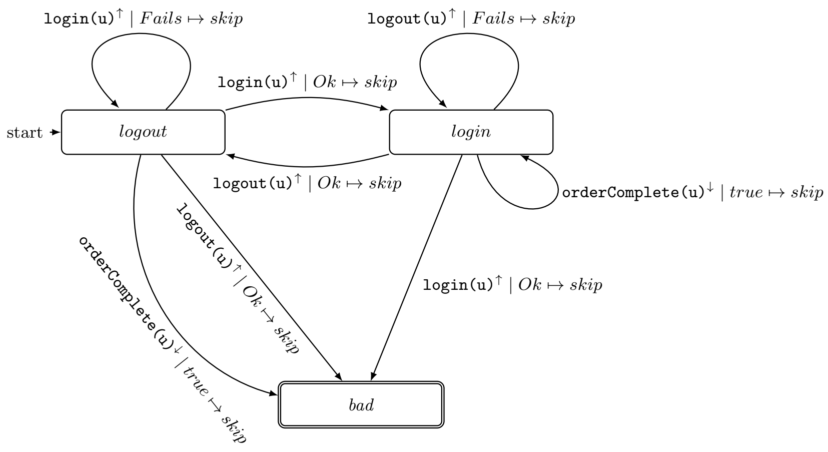


Fig. 15. *ppDATE* template describing properties about the log in and log out of users.

on the figures depicting the *ppDATE*s. Moreover, as the application is placed in a server, the monitor generated by our tool is placed in the server as well.

Login — Logout Users can freely browse through the web site of the application. However, if they want to buy products (i.e., pictures), they have to be logged in the application, to be able to proceed to the checkout section.

Fig. 14 and Fig. 15 illustrate the specification. The *ppDATE* in Fig. 14 creates instances of the *ppDATE* template *login-logout* whenever an object of class `User` is created, and the *ppDATE* template *login-logout* (Fig. 15) describes the following properties:

(i) *A user has to be logged in the application in order to perform a purchase, i.e., the checkout of a purchase can only happen between a login and a logout.*
(ii) *If a user is logged in, then that user cannot successfully log in again in the application until she logs out from it.*
(iii) *If a user is not logged-in, then that user cannot successfully log out from the application.*
(iv) *A user can only proceed to the checkout section if her status is a valid one.*
(v) *A user who is not a costumer cannot proceed to the checkout section.*

The transitions of the *ppDATE* described by the template control properties (i)–(iii). Initially, this *ppDATE* is in state *logout*. Then, whenever there is a successful login, the *ppDATE* moves to state *login*. Later, once the user logs out, the *ppDATE* returns to state *logout*. Therefore, if a purchase is performed (i.e., an order is checkout) while the *ppDATE* is in state *login*, then the *ppDATE* remains in that state. However, if a purchase is performed while the *ppDATE* is in state *logout*, then it shifts to state *bad*.[12] In addition, while being at state *logout*, if an attempt to log in is not successful, then the *ppDATE* stays in that state; and if there is a successful logout, then the *ppDATE* shifts to state *bad* due to the fact the user is considered to be logged out while the *ppDATE* is in that particular state. Something similar happens when the *ppDATE* is in state *login*. (In Fig. 15, *Fails* and *Ok* are abbreviations, for presentation purpose, of real Java expression checking the failure or success of the respective operations.)

Regarding properties (iv) and (v), they are addressed using Hoare triples. For instance, property (iv) is represented as follows:

```
{ !baseForm.getUserStatus().equals("Registered")
  && !baseForm.getUserStatus().equals("Unapproved"); }
prepareCheckout(baseForm)
{ \result.equals("success"); }
```

As the only non valid statuses are "Registered" and "Unapproved", if the status of the user is not one of these values, then starting a purchase, i.e., using method `prepareCheckout`, should return "success". Regarding property (v), a user is only considered to be a costumer if she has logged-in into the application. Even though this property seems to be similar to property (i), this similarity is only apparent. Property (i) only addresses the proper order in which the methods should be executed, whereas property (v) focuses on controlling how the data related to a user is modified during such executions. Finally, both properties (iv) and (v) are only placed in state *login* because that is the only state in which a successful purchase can occur, i.e., (iv) and (v) are context dependent data-oriented properties.

[12] Shifting to state *bad* means that a property was violated.

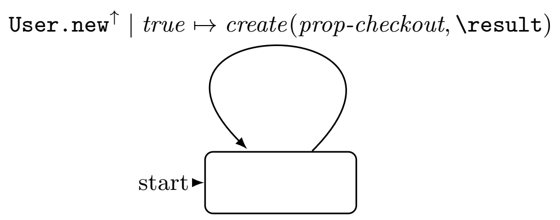

User.new$^{\uparrow}$ | *true* $\mapsto$ *create*(*prop-checkout*, \result)
start

Fig. 16. *ppDATE* in charge of creating instances of the template *prop-checkout*.

Purchases Checkout We consider that a purchase starts whenever an item (i.e., a product) is added to the cart. A user can continue either by adding other items to the cart or by removing some of the items from the cart. We refer to all the items in a cart as the *order*.

Once the user finishes the creation of her order, she may proceed to the checkout page. In SoftSlate, a checkout is realised in four steps. First, the user enters the contact information and delivery address. Then, the shipping method is selected (either ground transport or air transport), after which the user enters her credit card details. Finally, a confirmation for the order is requested. If accepted, the order is settled. Later, when the user receives the items, the order is considered to be completed.

Note that a user can modify her order as long as she has not yet confirmed it. If so, whenever she proceeds to the checkout section again, all its required steps have to be performed one more time. In addition, if the user removes all the items in an order, clears the cart or logs out[13], then the order is considered to be removed.

Fig. 16 and Fig. 17 illustrate a *ppDATE* specification where the *ppDATE* in Fig. 16 creates instances of the *ppDATE* template *prop-checkout* whenever an object of class `User` is created, and the *ppDATE* template *prop-checkout* (Fig. 17) describes the following properties:

(1) *The checkout of a purchase should be performed following the four required steps.*
(2) *It should not be possible to buy zero or less items.*
(3) *The expiration date of the credit card should not earlier than the current date.*
(4) *The price of a product should be positive.*
(5) *Before a purchase is completed, taxes should be processed.*
(6) *The total cost of a purchase should be equal to the sum of the prices of all the products to be purchased.*
(7) *If the price of an item changes, then its price in the order of the user should be updated.*

[13] Logging out clears the cart.

prop-checkout = $\lambda\ u$: `User`.

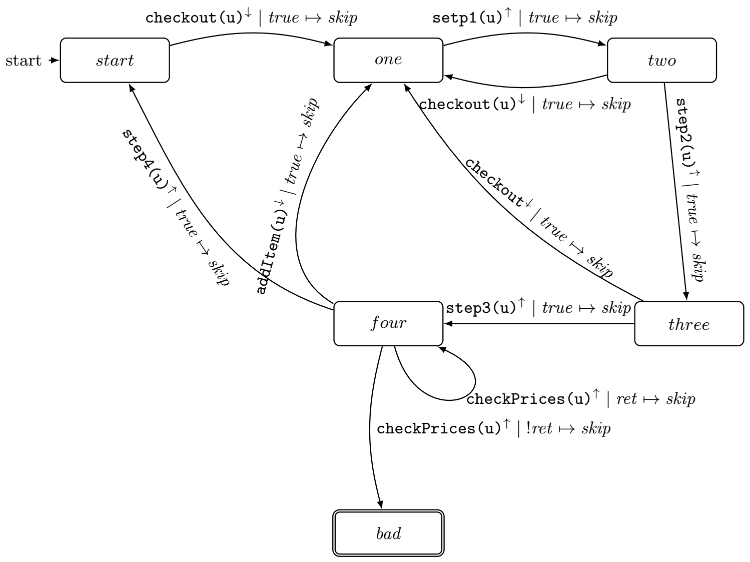


Fig. 17. *ppDATE* template describing properties related to checkout of purchases.

Again, consider the transitions of the *ppDATE* described by the template. When the first item is added to the cart, the *ppDATE* shifts to state *one*. In this state, once the first step of the checkout is completed, the *ppDATE* shifts to state *two*, and so on until reaching state *four*. In state *four*, once the order is settled, the *ppDATE* shifts back to state *start* in order to wait for a possible new purchase. Moreover, while being at either state *one*, *two*, *three* or *four*, if there is any change in the order, then the *ppDATE* shifts to state *one*, meaning that all the steps of the checkout have to be performed again. This is enough to control property (1).

Note that for readability reasons, in states *one*, *two*, *three* and *four* we have not included transitions going to state *start* whenever the user logs out, the cart is cleared or all the items in the cart are removed. In addition, we have not included transitions going to state *bad* from either state *one*, *two*, *three* or *four* if a step of the checkout was performed in a wrong way. For instance, if while being at state *one* either a second step, a third step or a fourth step of a purchase occurs instead of the first step, then the *ppDATE* shifts to state *bad*.

Regarding property (7), since the method in charge of updating the orders whenever the price of an item changes in the database is fully implemented using different Java libraries, writing an appropriate Hoare triple for it would require introducing several work-arounds. Instead, we implemented a method which compares the prices of the items in the order with their prices in the database, and include it as part of the information validation process corresponding to the fourth step of the purchase. Thereby, in state *four* there are two transitions controlling the result of this method. (Most real world applications of this kind would guarantee prices for some defined duration, and adjust it when that time has passed. For simplicity, we only model the latter in (7).)

Properties (2)–(6) are addressed with Hoare triples. Properties (2)–(4) are related to the integrity of the information introduced by either the users, in the case of (2) and (3), or the administrator, in the case of (4), on their requests to the server. Property (5) is related to the proper processing of taxes associated to the items in the current order. Property (6) enforces that the total amount that the user has to pay for her order should be equal to the sum of the totals of all the items included in the order.

As items could be added to the cart at any time during a purchase, property (2) is included in all the states of the *ppDATE*, with exception of the state *bad*.

On the other hand, property (3) is context dependent. This property should only be enforced on state *three*, which represents the step of a purchase where a user enters her credit card details. Note that, as it is in this case, a single property might be associated to several Hoare triples. For instance, below we introduce two of the four Hoare triples which describe property (3),

```
{ cardYear > actualYear; }
checkDate(cardMonth,cardYear, actualMonth,actualYear)
{ \result; }

{ cardYear < actualYear; }
checkDate(cardMonth,cardYear, actualMonth,actualYear)
{ !\result; }
```

Regarding property (4), we assume that initially all the data in the database is properly set. Therefore, this property should only be enforced every time that the administrator modifies the price of an item. As this may happen at any time during a purchase, this property is included in all the states of the *ppDATE*, with exception of the state *bad*.

In relation to property (5), in SoftSlate whenever the taxes of items are processed, the status of the order changes to “Tax processed”. This change is done by using the following method,

```
public void setStatus(String s) { status = s;}
```

This method might be simply specified as follows:

```
{ true; }  setStatus(s)  { status.equals(s); }
```

However, due to the fact that taxes are processed while the *ppDATE* is in state *four*, that we know which particular value should be written when updating the status of the order, i.e., “Tax processed”, and that *ppDATE* allows us to write context dependent properties, we include in *four* the following Hoare triple:

```
{ true; }  setStatus(s)  { status.equals("Tax␣processed"); }
```

Regarding property (6), it is represented by the following Hoare triple:

```
{ true; }
updateOrderAndDeliveryTotals(user,order,item)
{ user.getOrder().getSubtotal().doubleValue() ==
    (\old(user).getOrder().getSubtotal().doubleValue()
      + item.getTotal().doubleValue());}
```

In short, the new total amount is equal to the old total amount plus the amount of the newly added item.

9.2 Using StaRVOOrS

Since SoftSlate uses many Java libraries, to perform static analysis on its source code it was necessary to generate stub files for some of these libraries in order to allow KeY to find information about their method declarations.

Login — Logout When feeding StaRVOOrS with this property and the source code of SoftSlate, it automatically generates a runtime monitored version of the application and a report which summarises the results obtained from the static analysis.

Regarding the result of the translation, it consisted of a *DATE* specification which looks exactly like the original *ppDATE* specification. The static analysis and instrumentation process takes 11 seconds on a PC Pentium Core i7, where most time is used by KeY to statically analyse the Hoare triples (approximately 7 seconds). By inspecting the report we notice that KeY successfully verified all the Hoare triples in the *ppDATE* specification. Thus, the refined *ppDATE* specification to be translated was already a *DATE*, .i.e, the translation process did not have add any new transitions to the specification.

Purchases Checkout When feeding StaRVOOrS with this property and the source code of SoftSlate, it automatically generates a runtime monitored version of the application and a report which summarises the results obtained from the static analysis. The static analysis and instrumentation process takes 23 seconds on a PC Pentium Core i7, where most time is used by KeY to statically analyse the Hoare triples (approximately 20 seconds). By inspecting the report we can see that properties (2) and (3) are fully proved, properties (4) and (5) are not proved, and that property (6) and (7) are partially proved.

Regarding property (7), as KeY does not have any information about the state of purchases, and this property is context dependent, obviously, it is not able to prove it. However, thanks to the use of StaRVOOrS we can include this property in an appropriate state of the *ppDATE*, fact which guaranties that whenever a purchase reaches such state, this property is going to be verified at runtime by the generated monitor.

Regarding property (6), the report shows that this property postcondition is going to be checked upon entering method `updateOrderAndDeliveryTotals` only if the condition `user.getOrder() != null` holds. Thereby, this property is refined by StaRVOOrS as follows:

```
{ user.getOrder() != null; }
updateOrderAndDeliveryTotals(user,order,item)
{ user.getOrder().getSubtotal().doubleValue() ==
    (\old(user).getOrder().getSubtotal().doubleValue()
      + item.getTotal().doubleValue());}
```

This refined version of property (6) is the one verified by the generated monitor at runtime.

Finally, the result of the translation consisted on one *DATE* to create instances of the obtained *DATE* template *prop-checkout* (the translation of its homonymous *ppDATE* template), and three generated *DATE* templates whose instances verify properties (4)–(6). Note that the instances of the generated *DATE* templates are created by actions on the transitions of the *DATE* template *prop-checkout*.

9.3 Experimentation

Properties Analysis

Login — Logout Although this property may appear to be simple, by verifying it we discovered unexpected behaviour in SoftSlate when a user logs in, performs a purchase, and logs out. In spite of the fact that the user was logged in the application, the monitor flagged a violation of property (iii). It turned out that after performing the purchase, SoftSlate replaced the object representing the logged-in user by a new one.

More concretely, the log file generated by the monitor showed that a new monitor, corresponding to a new instance of the template *login-logout*, was generated for the 'new' user. So, we got two different user objects, the one who originally logged in into the system (let's call it u_{logged}) and the new generated one (let's call it u_{new}). The new monitor (corresponding to the user u_{new}) would then be in its initial state, that is in the state *logout*. Thus, when the (real) user tried to log out, the monitor corresponding to user u_{new} shifted to a *bad* state, while the monitor corresponding to *login-logout* = $\lambda\ u$: `User`.

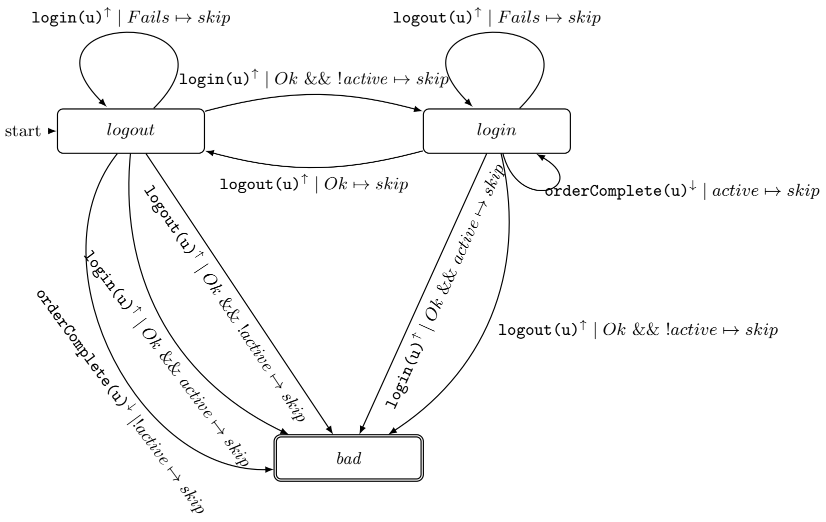

login(u)↑ | Fails ↦ skip
logout(u)↑ | Fails ↦ skip
login(u)↑ | Ok && !active ↦ skip
start
logout
login
logout(u)↑ | Ok ↦ skip
orderComplete(u)↓ | active ↦ skip
logout(u)↑ | Ok && !active ↦ skip
login(u)↑ | Ok && active ↦ skip
orderComplete(u)↓ |!active ↦ skip
logout(u)↑ | Ok && !active ↦ skip
login(u)↑ | Ok && active ↦ skip
bad

Fig. 18. Extension on the *ppDATE* describing properties related to the log in and log out of users illustrated in Fig. 15.

user u_{logged} remained in state *login*. As a consequence, property (iii) was violated.

In order to understand whether this is an error in the implementation we inspected the source code to better understand how the login and purchase were implemented. We found that each instance of class `User` was associated to a session, whose information was unique for each different execution of the application. Though the relation between (real) users and the session is bijective (for each real user there is a unique session, and vice versa), there were (at least) two instances of the class `User`, u_{logged} and u_{new}, associated with each session.

We were not sure what were the real reasons behind this design decision, but the implementation seemed correct, and our specification did not capture this situation. So, we decided to change our *ppDATE* template to capture this by including a Boolean variable reflecting whether the (real) user was connected or not, which we refer to as *active*. The updated *ppDATE* template is shown in Fig. 18. Further executions of the system (reproducing the previous executions and providing new ones) did not violate this property.

Purchases Checkout We also run the system many times in order to analyse whether the execution of SoftSlate fulfils the properties described by the provided *ppDATE* specification.

First, we performed several purchases to analyse if property (1) was fulfilled. We added some items to the cart, bought them, and added and removed items at any stage of the checkout of a purchase, and then completed the purchase. None of these operations violated this property. We re-run the system executing the same steps as above to check property (5), which was not violated.

Next, we continued performing purchases, but this time the administrator of the application introduced modifications in the price of some items during the purchases. By doing so we were able to analyse whether properties (4), (6) and (7) were violated.[14]

In order to check whether property (4) held, we executed the system logged in as administrator and as a normal user (in parallel). The user performed a purchase (and thus the item was added to the cart), and as administrator we modified the price of the item introducing a negative value as its new price. At this moment the monitor reported that property (4) was violated. By inspecting the price of the modified item in the database, we could confirm that the negative value provided by the administrator was actually assigned to the item. This clearly was an error. We corrected this by not allowing to input negative numbers, and thus property (4) was finally satisfied.

On the other hand, when the administrator modified the price of an item introducing a positive value as its new price, then property (4) was fulfilled as expected. However, we noticed that property (7) was violated: some of the prices of the items in the order did not match with the prices in the database.[15] In particular, the mismatched values were those that were modified by the administrator: the new prices were propagated to the database but they were not updated in the visualisation of the cart (to the user). This was an error, and when inspecting the code we realised that there was a method implementing the propagation of the update, but it was not called when the change (done by the administrator) was performed. We have not yet corrected this error in the original code.

Property (6) was not violated by any of the previous executions.

Runtime Verification Overhead Analysis In this section we analyse the overhead added to SoftSlate by the monitor generated using StaRVOOrS. To perform this analysis, we considered three scenarios: several users performed one purchase, 10 purchases in a row, and 100 purchases in a row.

Table 1 shows the average execution time of: (a) an unmonitored execution of SoftSlate; (b) a monitored execution of SoftSlate using the

[14] Remember that properties (2) and (3) were fully proved statically.

[15] This also happened when entering negative numbers, but we only found out this when focusing on checking property (7) after correcting the issue with negative inputs.

Purchases	(a) no monitoring	(b) monitoring S	(c) monitoring S'
1	800 ms	1,300 ms	1,100 ms
10	10,500 ms	15,500 ms	13,000 ms
100	120,000 ms	190,000 ms	150,000 ms

Table 1. Performance of different purchases.

original *ppDATE* specification S, and (c) a monitored execution of SoftSlate using specification S', obtained from S via static (partial) proof analysis using StaRVOOrS. In all three scenarios, the users and the server hosting SoftSlate were run in different computers with identical specifications (a PC Intel Core i7 using a single core). Note that as SoftSlate is an interactive application, in order to perform these experiments we have implemented a program which uses url connections to access the application and perform a purchase[16]. Therefore, our experiments consist on executing this program repeatedly and measuring its execution time.

As expected, adding a monitor to SoftSlate introduced overhead on its execution time. However, when we compared the overhead added by the monitor which uses the original *ppDATE* specification (without optimisations) (b), with the one added by the monitor which was generated using StaRVOOrS (c), one could notice a reduction in overheads gained by using our tool.

Through optimisations introduced by StaRVOOrS, we obtained a version of the monitor which, in relation to the times in (a), introduced in average a 25% of overhead to the execution time of the system. On the contrary, the monitor without the optimisations of StaRVOOrS introduced a 50% of overhead to the execution time.

Even though these results are not as impressive as the one we obtained on the case study analysed in [6] (Mondex, also reported here in Sec. 10), the monitor generated by our tool for SoftSlate still has a better performance than the one which uses the original *ppDATE* specification. The main difference lies in the amount of Hoare triples which have to be runtime verified in each case study. Every time an experiment is performed to analyse SoftSlate, the optimise monitor generated by StaRVOOrS verifies 3 Hoare triples, whereas the monitor using the original *ppDATE* specification (without optimisations) verifies 5. However, each experiment performed on Mondex requires the verification of 7 Hoare triples when using the unoptimised version of the monitor, whereas the optimised one does not have to verify any Hoare triples at all (cf. Sec. 10).

[16] The package java.net is used here to handle the communication between our program and SoftSlate.

10 Case study: Mondex

Mondex is an electronic purse application which is used by smart cards products [2], and has been considered as a verification benchmark problem since 2006, originally appearing as case study as part of the Verified Software Grand Challenge [36]. Mondex's original sanitised specification can be found in [32]. It consists of a Z specification [31], together with hand-written proofs of several properties.

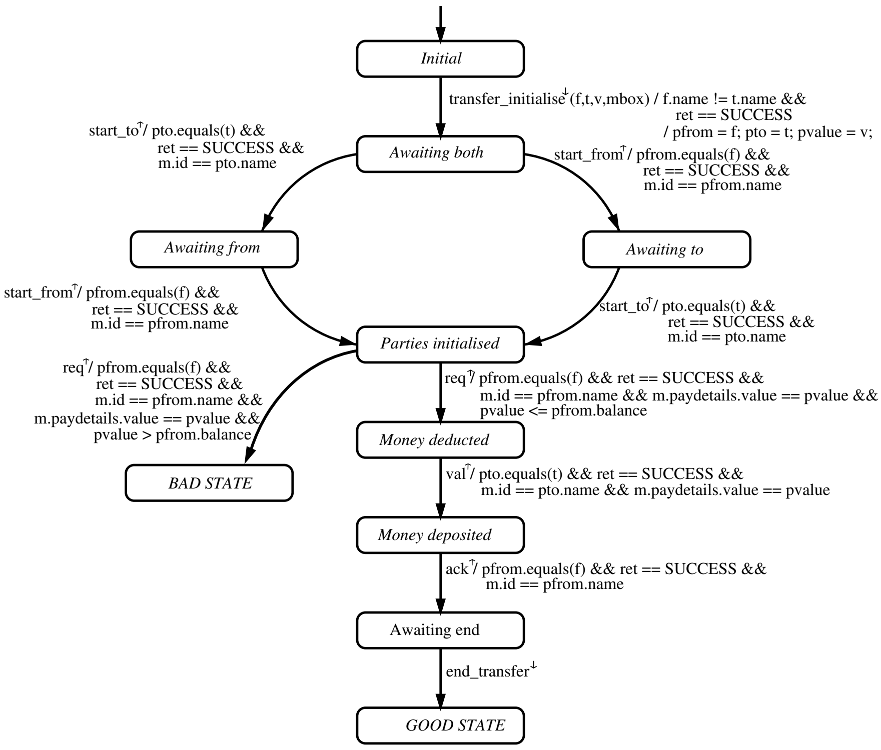
Mondex essentially provides a financial transaction system supporting transferring of funds between accounts, or *purses*. Whenever a person has to make a transaction, electronic money is taken from their electronic purse and transferred to the target electronic purse. Such transactions are performed following a multi-step message exchange protocol: (1) the source and destination purses should (independently) register with the central fund transferring manager; (2) await a request to deduct funds from the source purse; (3) await a request to add the funds to the destination purse; and finally (4) an acknowledgement is sent to indicate that the transfer took place before the transaction ends.

In our version of this case study we consider a Java implementation running on a desktop computer instead of a Java Card implementation running on smart cards. The principal difference in the implementation is that in our version some methods return values to indicate whether their output is normal or erroneous, instead of raising Java Card exceptions. Our specification is strongly inspired by the JML formalisation presented in [34]. The full specification and source code of our case study can be found in [4]. The specification (see Fig. 19) consists of a *ppDATE* with 10 states, 25 transitions and a total of 26 different Hoare triples. The implementation consists on 514 lines of code (without comments) which are distributed over 8 files.

Note that *ppDATE* allows us to represent the overall status of the observer using *ppDATE* states. In other pre/post-style specification approaches, one would instead introduce additional data, and corresponding additional constraints, as is indeed done in [34] when specifying Mondex with JML. Such additional data implies a certain complexity of the specification, which somehow lacks the structure of the problem. We believe that specifications of this kind are sometimes developed with an automaton in mind. In *ppDATE*, we can make that automaton explicit. This being said, we want to stress again that we took great advantage of the JML specification of Mondex in [34].

10.1 *ppDATE* Property

Fig. 19 illustrates a *ppDATE* describing the top-level specification of Mondex. To keep the *ppDATE* readable, the description of the different


Initial
transfer_initialise$^{\downarrow}$(f,t,v,mbox) / f.name != t.name && ret == SUCCESS / pfrom = f; pto = t; pvalue = v;
Awaiting both
start_to$^{\uparrow}$/ pto.equals(t) && ret == SUCCESS && m.id == pto.name
start_from$^{\uparrow}$/ pfrom.equals(f) && ret == SUCCESS && m.id == pfrom.name
Awaiting from
Awaiting to
start_from$^{\uparrow}$/ pfrom.equals(f) && ret == SUCCESS && m.id == pfrom.name
start_to$^{\uparrow}$/ pto.equals(t) && ret == SUCCESS && m.id == pto.name
Parties initialised
req$^{\uparrow}$/ pfrom.equals(f) && ret == SUCCESS && m.id == pfrom.name && m.paydetails.value == pvalue && pvalue > pfrom.balance
req$^{\uparrow}$/ pfrom.equals(f) && ret == SUCCESS && m.id == pfrom.name && m.paydetails.value == pvalue && pvalue <= pfrom.balance
BAD STATE
Money deducted
val$^{\uparrow}$/ pto.equals(t) && ret == SUCCESS && m.id == pto.name && m.paydetails.value == pvalue
Money deposited
ack$^{\uparrow}$/ pfrom.equals(f) && ret == SUCCESS && m.id == pfrom.name
Awaiting end
end_transfer$^{\downarrow}$
GOOD STATE

In addition:

* All states have outgoing transitions for ret == SUCCESS && SENDER!=party (where party is the party from whom a message is not expected) going to a bad state.
* All states but 'Awaiting end' have outgoing transitions for end_transfer, going to a bad state.

Fig. 19. *ppDATE* to monitor the behaviour of the transaction protocol

Hoare triples are not included in the figure. (We will show some of them below.)

At the automaton level, the *ppDATE* specifies the control-oriented property which indicates how the multi-step message exchange protocol is suppose to work. For instance, after the parties are initialised (encoded in state Parties Initialised), a message requesting to transfer more money than the one available in the source purse should fail. Otherwise, such a message should take the *ppDATE* to a state in which the protocol now allows for the money to be transferred to the destination purse (named Money deducted). Note that the *ppDATE* will not take any explicit action whenever the state BAD STATE is reached. It will stay in this state until the whole monitor is restarted.

In contrast, the pre/postconditions properties placed on the states of the *ppDATE* ensure the well-behaviour of the methods involved in the individual steps of the protocol, behaviour which obviously changes together with the status of the protocol. For instance, once two purses agree on participating in a money transfer and the destination purse has requested for certain amount of money, (encoded in state Money Deducted), method `val_operation` which transfers money from the source purse to the destination one should succeed and increase the money of the destination purse by the sent amount (provided the limit of its account has not been reached), as shown in the Hoare triple below:

```
{ checkSameTransaction() == SUCCESS
  && transaction.value <= (ShortMaxValue - balance); }
val_operation
{ \result == SUCCESS
  && (balance == \old(balance) + transaction.value); }
```

On the other hand, if the same method is accessed after the funds have already been transferred (encoded in state Money deposit), then the destination purse content should remain unchanged, and the request should be ignored:

```
{ checkSameTransaction() == SUCCESS
  && transaction.value <= (ShortMaxValue - balance); }
val_operation
{ \result == IGNORED; }
```

Note that both Hoare triples above have the same precondition, but depending on the state of the *ppDATE* (i.e., the state of the protocol) different behaviours (i.e., postconditions) are expected for method `val_operation`.

10.2 Using StaRVOOrS

Running StaRVOOrS on the source code of Mondex and the *ppDATE* depicted in Fig. 19 automatically produces a runtime monitored version of the application and a report summarising the results obtained from the static analysis. The static analysis and instrumentation process takes 1 minute 20 seconds on a PC Pentium Core i7, where most time is used by KeY to statically analyse the Hoare triples (approximately 1 minute 15 seconds).

The monitor generated by our tool consists one *DATE* to control the main property, and 24 *DATE*s templates to control the postconditions which were only partially verified by KeY, with 106 states and 196 transitions in total. By inspecting the report we can see that the two Hoare triples associated to the initialisation and termination of a transaction were fully proved, and that all the other 24 triples about the methods involved in the transaction protocol were the partially verified ones. For instance, let us consider the property already discussed in the previous section about method `val_operation`, which we will refer here to as *val_operation_ok*:

```
{ checkSameTransaction() == SUCCESS
  && transaction.value <= (ShortMaxValue - balance); }
val_operation
{ \result == SUCCESS
  && (balance == \old(balance) + transaction.value); }
```

The report shows that the postcondition will have to be checked at runtime only when the condition `status != 2` holds upon entering `val_operation` (i.e., the destination purse is not waiting for the arrival of the requested money). Thus, the previous Hoare triple was refined by StaRVOOrS as follows:

```
{ checkSameTransaction() == SUCCESS
  && transaction.value <= (ShortMaxValue - balance)
  && !(status == ProtocolStatus.Epv); }
val_operation
{ \result == SUCCESS
  && (balance == \old(balance) + transaction.value); }
```

This refined version of the property is the one which will be runtime verified by the generated monitor.

The size of the source code of the original implementation of Mondex was 23.5kB. After running the tool, the total size of all the generated files (i.e. instrumented version of the source code and the implementation of the monitor) grows to 277.4kB.

10.3 Experimentation

We now summarise the experimental results of applying our approach to the Mondex case study.

Normal Behaviour The table illustrated in Fig. 20 shows the execution time of: (a) an unmonitored implementation of Mondex; (b) a monitored implementation using the original *ppDATE* specification S, and (c) a monitored implementation using specification S', obtained from S via static (partial) proof analysis using StaRVOOrS. In all three scenarios, the system is run over a numbers of transactions which do not violate the specification. Note that in case (c), statically analysing all the Hoare triples took KeY around 1 minute, which however is done once and for all prior to deployment. These scenarios were analysed on a PC Intel Core i7 using a single core.

As one would expect, the addition of a monitor to the system introduces execution time overhead (b). However, if we compare this overhead to the one added by the monitor which was generated by StaRVOOrS (c), one can see a substantial overhead reduction, gained through the use of our tool. Through our optimisations we obtain a version which is at least 10 times faster for a low number of transactions, and this factor rises up to 900 when the number of transactions is increased. This significant

Transactions	(a) no monitoring	(b) monitoring S	(c) monitoring S'
10	8 ms	120 ms	15 ms
100	50 ms	3500 ms	90 ms
1000	250 ms	330000 ms	375 ms

Fig. 20. Performance of different transactions which do not violate any of the specified properties

reduction in execution time overheads is mainly due to the fact that monitoring data-centric properties may be prohibitively expensive. In fact, using S, each method invocation involved in the transfer protocol creates an additional *DATE* that will check the postcondition on exit. However, the postcondition checker is only created if the precondition holds on method invocation. In this case study, this causes large overheads when monitoring the unoptimised specification. Using the results from static verification, however, strengthens the preconditions by additional constraints, which in the Mondex case state were always falsified at invocation time, meaning that no postcondition checker is ever created. Apparently, in Mondex, the algorithmic complexity of the individual method implementations is limited enough such that KeY could fully prove the methods correct (automatically) *if only* the internal constraints corresponding to the *ppDATE* states were provided to KeY. But as they are not, KeY generates those constraints (closed branch conditions, see Sec. 4), and adds their negation to the preconditions. With that, the preconditions are never true at runtime. This phenomenon cannot be fully generalised to cases where KeY really lacks (automated) proving power for the code at hand, or where the code is faulty of course.

Faulty Behaviour Usually, it is hard to get full proofs when using a static verifier like KeY without considering either user interaction with the prover or the use of special annotations, e.g., loop invariants, to help the prover on its task. However, it might be the case that the static verifier does not succeed in closing a branch in the proof due to the fact that the remaining open goal was generated by an erroneous execution path. KeY cannot *per se* determine which one of these situations is dealing with. Fortunately, Larva can detect the occurrence of the erroneous case whenever it appears at runtime.

We have intentionally injected errors into Mondex source to verify that the optimised monitor still detects them. Consider the case of a bug in the implementation of method `val_operation` — the value of variable `balance` is incremented with a different amount from the one given in the specification of the method. When analysing property `val_operation_ok`, KeY obviously does not manage to prove it. Therefore, the whole property will have to be runtime verified. The monitor spots this error reaching a bad state

In addition, we have also considered incomplete and wrong specifications. In the case where the specification is too weak, the implementation may fulfil it for wrong reasons. As in all verification approaches, we may not catch this kind of problem. When using our verification approach there lies the possibility that the problem propagates to a state in which the specification is strong enough to identify it. For example, consider if the specification does not specify how the variables of a purse should be initialised by the `ConPurse` class constructor, and there is an implementation error where the variable `balance` is initialised to -1 instead of being initialised to 0. In spite of the error in the specification, KeY would proceed normally with the proofs and the previous particular situation would not be directly controlled on runtime. However, this erroneous initialisation leads to an erroneous initial charge of money in the purses (performed using the method `chargeMoney` in class `ConPurse`). As `balance` is negative, the previous method fails to update it with the new amount of money. Hence, after applying `chargeMoney` the value of `balance` is still -1. Thereby, whenever a purse tries to begin a transfer, either the method initialising the sender purse during a transaction or the method initialising the receiver purse during a transaction will fail its execution (the former due to insufficient funds and the latter due to a value overflow). This failure leads to an unsuccessful termination of the transfer, which is detected by the monitor controlling the transaction protocol and takes it to a bad state. This analysis can be easily conclude by inspecting the execution trace generated by the monitor. This trace allows one to backtrack through the execution of the different methods until reaching which was the problem which was the cause the failure. In this scenario, it is important to note that in spite of the fact that we have not enforced any Hoare triple on the constructor of class `ConPurse`, it was specified and proved correct using KeY.

On the other hand, if a Hoare triple has an overly weak precondition or overly strong postcondition, then KeY will fail to prove the Hoare triple. StaRVOOrS thus ensures that the Hoare triple is checked at runtime, which allows us to realise when expected results arise. Finally, another scenario is when the user uses erroneous data, not detected by the application. For instance, a user might request a transfer exceeding the amount of money in a purse. In this situation, the method initialising the sender purse during a transaction will fail its execution due to insufficient funds and this will lead to an unsuccessful termination of the transfer. This unsuccessful termination is detected by the runtime monitor controlling the transaction protocol.

11 Related Work

The combination of different verification techniques is gaining more and more popularity. One active area of research is the combination of testing and static analysis, e.g. [11, 16, 18, 21, 25, 26, 33]. A direct comparison of our work with those would not be fully fair as we have different objectives. We are not aiming at generating test cases, but at monitoring the actual post-deployment runs of the system. What we have in common is that static analysis/verification is used to limit the dynamic efforts, there by filtering test cases, here by filtering checks at runtime.

Another line of research is the combination of testing and runtime verification. Decker *et al.* in [23] introduce an extension of the testing framework JUnit, which adds runtime verification artefacts to it. In this extension, during the execution of a test, a monitor is in charge of checking whether the actual executed test conforms with the property being monitored. In [10] Artho *et al.* present a framework where automated test case generation benefits from the use of runtime verification in a similar way to [23]. Falzon and Pace [24] study the combination of QuickCheck and Larva by presenting a technique which extracts monitors from a QuickCheck testing specifications. Even though this line of work have a different objective compare to ours, it is worth mentioning that the QuickCheck automata used in [24] are quite similar to *ppDATE*s. QuickCheck automata employ pre/postconditions as part of their transitions, as opposed to *ppDATE*s which include them in the states of the automata. This similarity may suggest that it might be possible to extend our approach by also including the possibility of perform testing.

Another area worth mentioning is the combination of runtime assertion checks with runtime verification. In [22] de Boer *et al.* present SAGA, a framework which combines runtime assertion checking with monitoring. In contrast to our approach which targets general data- and control-oriented properties, SAGA focuses on the verification of both data-flow and control-flow properties of Java classes and interfaces, e.g., interaction protocol among objects. However, we are mainly interested in the combination of static verification and runtime verification such that static verification is used to reduce the overhead introduced to the system execution by monitoring properties. Wonisch *et al.* in [35] make use of program transformations in order to avoid unsafe program executions. In [14] the efficiency of runtime monitoring based on tracematches is improved by using a static analysis technique which reduces the runtime instrumentation needed. The technique consists on three stages: exclusion of some tracematches, elimination of inconsistent instrumentation points, and additionally refinement of this analysis considering the order of execution.

Other works use this kind of combination but with different goals. In [15] Bodden and Lam present CLARA, a framework which uses static techniques aiming to improve the monitors themselves, instead of verifying software. The work by Zee *et al.* in [37] investigates the combination of static and runtime verification, but aiming at a specification language whose specifications may be both statically and runtime checked. With this goal in mind, they extend the static verifier Jahob by adding techniques to verify specifications at runtime. In this approach, most of the properties which can be verified are data-oriented, as opposed to ours where control-oriented properties are covered as well. In [30] Sözer integrates static code analysis and runtime verification. On this approach, runtime verification statements are created from static code analysis alerts, in order to generate monitors which will allow to both check for possible faults in the system and eliminate false positives obtained in the static phase.

Many specification approaches, such as SPARK [12], JML [28] and SPEC# [13] are supported by both static and runtime verification tools. Nevertheless, to the best of our knowledge, static verification is not used to optimise the runtime verification of properties.

12 Conclusions

In this paper we have presented StaRVOOrS, a framework for verifying integrated data- and control-oriented properties for Java programs, using a combination of static and runtime verification. The StaRVOOrS tool-chain uses KeY [5] for static verification, and Larva [20] for the verification performed at runtime.

We have presented the language *ppDATE* which is based on automata and pre/post conditions to describe properties of both, the control flow and the data computations. The basic structuring principle of the language is the composition of parallel automata, whose transitions fire simultaneously in reaction to events of the observed system, but also in reaction to events generated by some automata in the previous step. A distinguishing feature of the language is the inclusion of functional properties of computation units into the above, thereby capturing the dependency of functional properties on the history of previous events, by assigning Hoare triples to (automata-theoretic) states. Finally, the template concept allows to parameterise components in a great variety of ways, and create concrete instantiations dynamically.

We also presented here a semantics of *ppDATE*s, precisely describing the interplay of transitions, event consumption and generation, Hoare triple monitoring, creation of template instances. We then use the semantics to prove soundness of the algorithm our tool uses to translate *ppDATE* into *DATE*, allowing us to employ the *DATE* tool Larva as a back-end for runtime verifying *ppDATE* specifications.

This article also reports on the application of StaRVOOrS to SoftSlate, an open-source shopping cart web application. In this case study, we analyse *ppDATE*s describing properties about the proper behaviour of the system while users perform purchases. We also report on application of STARVOORS to the verification benchmark Mondex, an electronic purse application. We demonstrate how properties can be verified using combined static and runtime verification.

For SoftSlate, the overhead of pure runtime verification (without employing static verification) is roughly 50%, a penalty which we get down to roughly 25% when using STARVOORS, by facilitating static verification (cf. Section 9.3). These differences are much smaller compared to when we applied STARVOORS to the Mondex case study, where pure runtime verification created a much higher overhead. Compared to that, the monitor created by STARVOORS was 10 times faster for a low number of transactions, and up to 900 times faster as the number of transactions increase. 'When using the monitor generated from the original specification provided for Mondex, the execution of each method involved in a transaction (7 in total) creates an additional *DATE* to be traversed in parallel, which is in charge of checking the postcondition. This would lead to the large overheads obtained in that case study. However, when using the monitor generated by STARVOORS, thanks to the optimisations introduced in the specification by this tool, no additional *DATE*s are created when a transaction is performed, because the additional checks in the preconditions are false at runtime.

As a final remark, note that the efficiency gain for monitoring will benefit from any improvements in the used static and runtime verifiers. For instance, if KeY is improved in such a way that more branches are closed during the static proof, then this will have an immediate effect in STARVOORS thus reducing the runtime overhead. Similarly, any optimisation performed in LARVA will only bring benefits to our tool.

We are currently looking at ways of pushing our techniques further. On one hand, we are looking at techniques to add control-flow static analysis to STARVOORS, thus benefiting from further optimisation prior to deployment. We are also looking at extending the framework to deal with distributed systems [9], which brings in new challenges, and might require assume-guarantee reasoning to enable us to perform static analysis based optimisations.

References

1. Apache Tomcat. `tomcat.apache.org/`.
2. MasterCard International Inc. Mondex. `www.mondexusa.com/`.
3. SoftSlate Commerce. `www.softslate.com/`.
4. StaRVOOrS web page. `www.cse.chalmers.se/~chimento/starvoors`.
5. Wolfgang Ahrendt, Bernhard Beckert, Richard Bubel, Reiner Hähnle, Peter H. Schmitt, and Mattias Ulbrich, editors. *Deductive Software Verification—The KeY Book*, volume 10001 of *LNCS*. Springer, 2016. to appear.
6. Wolfgang Ahrendt, Jesús Mauricio Chimento, Gordon J. Pace, and Gerardo Schneider. A Specification Language for Static and Runtime Verification of Data and Control Properties. In *FM'15*, volume 9109 of *LNCS*. Springer, 2015.
7. Wolfgang Ahrendt and Maximilian Dylla. A system for compositional verification of asynchronous objects. *Science of Computer Programming*, 2012.
8. Wolfgang Ahrendt, Gordon Pace, and Gerardo Schneider. A Unified Approach for Static and Runtime Verification: Framework and Applications. In *ISoLA'12*, LNCS 7609. Springer, 2012.
9. Wolfgang Ahrendt, Gordon J. Pace, and Gerardo Schneider. StaRVOOrS - Episode II - Strengthen and Distribute the Force. In *ISoLA'16 (1)*, LNCS 9952. Springer, 2016.
10. C. Artho, H. Barringer, A. Goldberg, K. Havelund, S. Khurshid, M. Lowry, C. Pasareanu, G. Rosu, K. Sen, W. Visser, et al. Combining test case generation and runtime verification. *Theoretical Computer Science*, 336(2-3):209–234, 2005.
11. Cyrille Artho and Armin Biere. Combined Static and Dynamic Analysis. In *AIOOL'05*, volume 131 of *ENTCS*, pages 3–14, 2005.
12. John Barnes. *SPARK: The Proven Approach to High Integrity Software*. Altran Praxis, http://www.altran.co.uk, UK, 2012.
13. Mike Barnett, K. Rustan M. Leino, and Wolfram Schulte. The Spec# programming system: An overview. In *CASSIS'05*, volume 3362 of *LNCS*, pages 49–69. Springer, 2005.
14. Eric Bodden, Laurie J. Hendren, and Ondrej Lhoták. A Staged Static Program Analysis to Improve the Performance of Runtime Monitoring. In *ECOOP'07*, LNCS 4609, 2007.
15. Eric Bodden and Patrick Lam. Clara: Partially Evaluating Runtime Monitors at Compile Time - Tutorial Supplement. In *RV'10*, volume 6418 of *LNCS*, pages 74–88, 2010.
16. Lilian Burdy, Yoonsik Cheon, David R. Cok, Michael D. Ernst, Joseph R. Kiniry, Gary T. Leavens, K. Rustan M. Leino, and Erik Poll. An overview of JML tools and applications. *International Journal on Software Tools for Technology Transfer*, 7(3):212–232, 2005.
17. Jesús Mauricio Chimento, Wolfgang Ahrendt, Gordon J. Pace, and Gerardo Schneider. StaRVOOrS: A Tool for Combined Static and Runtime Verification of Java. In Ezio Bartocci and Rupak Majumdar, editors, *Runtime Verification*, volume 9333 of *Lecture Notes in Computer Science*, pages 297–305. Springer International Publishing, 2015.

18. Maria Christakis, Peter Müller, and Valentin Wüstholz. Collaborative verification and testing with explicit assumptions. In *FM'12: Formal Methods - 18th International Symposium, Paris, France, August 27-31, 2012. Proceedings*, pages 132–146, 2012.
19. Christian Colombo, Gordon J. Pace, and Gerardo Schneider. Dynamic Event-Based Runtime Monitoring of Real-Time and Contextual Properties. In *FMICS'08*, volume 5596 of *LNCS*, pages 135–149. Springer-Verlag, September 2009.
20. Christian Colombo, Gordon J. Pace, and Gerardo Schneider. LARVA - A Tool for Runtime Monitoring of Java Programs. In *SEFM'09*, pages 33–37. IEEE Computer Society, 2009.
21. Christoph Csallner and Yannis Smaragdakis. Check 'n' crash: combining static checking and testing. In *27th International Conference on Software Engineering (ICSE 2005), 15-21 May 2005, St. Louis, Missouri, USA*, pages 422–431, 2005.
22. Frank S. de Boer, Stijn de Gouw, Einar Broch Johnsen, and Peter Y. H. Wong. Run-time checking of data- and protocol-oriented properties of Java programs: an industrial case study. In Sung Y. Shin and Jos Carlos Maldonado, editors, *SAC*, pages 1573–1578. ACM, 2013.
23. Normann Decker, Martin Leucker, and Daniel Thoma. jUnitRV - Adding Runtime Verification to jUnit. In *NASA Formal Methods*, volume LNCS 7871. Springer-Verlag Berlin Heidelberg, Springer-Verlag Berlin Heidelberg, 2013.
24. Kevin Falzon and Gordon Pace. Combining Testing and Runtime Verification Techniques. In *Model-based Methodologies for Pervasive and Embedded Software*, volume LNCS 7706, 2012.
25. Cormac Flanagan, K. Rustan M Leino, Mark Lillibridge, Greg Nelson, James B. Saxe, and Raymie Stata. Extended Static Checking for Java. In Jens Knoop and Laurie J. Hendren, editors, *PLDI'02*, pages 234–245. ACM, 2002.
26. Xi Ge, Kunal Taneja, Tao Xie, and Nikolai Tillmann. DyTa: Dynamic Symbolic Execution Guided With Static Verification Results. In *Proceedings of the 33rd International Conference on Software Engineering, ICSE 2011, Waikiki, Honolulu , HI, USA, May 21-28, 2011*, pages 992–994, 2011.
27. David Gries. *The Science of Programming.* Springer-Verlag New York, Inc., Secaucus, NJ, USA, 1st edition, 1987.
28. Gary T. Leavens, Erik Poll, Curtis Clifton, Yoonsik Cheon, Clyde Ruby, David Cok, Peter Müller, Joseph Kiniry, and Patrice Chalin. *JML Reference Manual. Draft 1.200*, 2007.
29. Florence Maraninchi and Yann Rémond. Running-modes of real-time systems: a case-study with mode-automata. In *12th Euromicro Conference on Real-Time Systems (ECRTS 2000), 19-21 June 2000, Stockholm, Sweden, Proceedings*, pages 257–264, 2000.
30. Hasan Sözer. Integrated static code analysis and runtime verification. *Softw., Pract. Exper.*, 45(10):1359–1373, 2015.
31. J. M. Spivey. *The Z Notation: A Reference Manual.* Prentice-Hall, Inc., Upper Saddle River, NJ, USA, 1989.
32. Susan Stepney, David Cooper, and Jim Woodcock. An Electronic Purse: Specification, Refinement and Proof. *Technical monograph PRG-126, Oxford University Computing Laboratory*, 2000.

33. Nikolai Tillmann and Jonathan de Halleux. Pex-White Box Test Generation for .NET. In Bernhard Beckert and Reiner Hhnle, editors, *TAP*, volume 4966 of *Lecture Notes in Computer Science*, pages 134–153. Springer, 2008.
34. Isabel Tonin. Verifying the Mondex Case Study. The KeY Approach. *Technical Report 2007-4, Universität Karlsruhe*, 2007.
35. Daniel Wonisch, Alexander Schremmer, and Heike Wehrheim. Zero Overhead Runtime Monitoring. In *SEFM'13*, volume 8137 of *LNCS*, pages 244–258. Springer Berlin Heidelberg, 2013.
36. Jim Woodcock. First Steps in the Verified Software Grand Challenge. In *SEW'06*, pages 203–206. IEEE Computer Society, 2006.
37. Karen Zee, Viktor Kuncak, Michael Taylor, and Martin C. Rinard. Runtime Checking for Program Verification. In *RV'07*, volume 4839 of *LNCS*, pages 202–213. Springer, 2007.

A Proofs of Coupling Invariant Lemmas

In order to prove both Lemma 2 and Lemma 3, we introduce the following two propositions. Prop. 1 says that the translation algorithm only modifies the actions of the transitions in the translated *ppDATE* network. Prop. 2 says that for every transition in the translation either there is a similar transition in the original *ppDATE* network, or there is not such a transition, due to the fact that the transition is a new loop transition (added by the translation to control Hoare triples).

Remember that we represent the translation of a single *ppDATE* to *DATE* with the function $\kappa \in ppDATE \mapsto DATE$.

Proposition 1. *Given a ppDATE network $pn = (M, V, \nu_0, T_{ppd})$ and its translation*
ppd2DATE(pn) $= (M', V, \nu_0, T'_d)$,

$$\begin{aligned}
&\forall\ m, q, q', tr, c, a\ \cdot \\
&\quad q \xrightarrow{tr|c\mapsto a}_{m} q' \ \textit{and}\ m \in M\ \textit{and}\ \kappa(m) \in M' \cdot \\
&\quad (\exists\ a' \cdot q \xrightarrow{tr|c\mapsto a'}_{\kappa(m)} q')
\end{aligned}$$

Proof. Given a *ppDATE* $m \in M$ and a state $q \in Q_m$, whenever $\Pi_m(q) = \emptyset$, $\Pi_m(q) \neq \emptyset$ but there is no Hoare triple associated to the method related to trigger tr, or the trigger is associated to exiting a method, by Step 3b., transitions remain unchanged in the translation. Therefore, $a' = a$ in these cases.

On the other hand, for each clashing Hoare triple $\{\pi\}\,\sigma\,\{\pi'\} \in \Pi_m(q)$, by step $3a_2$., the transition $q \xrightarrow{tr|c\mapsto a}_{m} q'$ is replaced by one of the following transitions:

$q \xrightarrow{tr|c\mapsto\{a;\,\texttt{if}\ \pi\ \texttt{then}\ \texttt{create}(post\_checker,(\sigma_{id}^{\uparrow},\pi'))\}}_{\kappa(m)} q'$, or

$q \xrightarrow{tr|c\mapsto\{a;\,\texttt{if}\ \pi\ \texttt{then}\ \texttt{create}(post\_checker\_h,(\sigma_{id}^{\uparrow},val_i))\}}_{\kappa(m)} q'$.

Thereby, either $a' = a;\ \texttt{if}\ \pi\ \texttt{then}\ \texttt{create}(post\_checker, (e_{id}^{\uparrow}, \pi'))\}$, or
$a' = a;\ \texttt{if}\ \pi\ \texttt{then}\ \texttt{create}(post\_checker\_h, (e_{id}^{\uparrow}, val_i))\}$.

Finally, as in step $3a_1$ non-clashing Hoare triples add new transitions but do not modified existing ones, this case trivially holds. □

Proposition 2. *Given a ppDATE network $pn = (M, V, \nu_0, T_{ppd})$ and its translation*
ppd2DATE(pn) $= (M', V, \nu_0, T'_d)$,

$$\begin{aligned}
&\forall\ m', q, q', tr, c, a\ \cdot \\
&\quad q \xrightarrow{tr|c\mapsto a}_{m'} q' \ \textit{and}\ m' \in M' \cdot \\
&\quad\quad (\exists\ m, a' \cdot m \in M, \kappa(m) = m' \cdot q \xrightarrow{tr|c\mapsto a'}_{m} q') \\
&\quad\quad \textit{or} \\
&\quad\quad ((\nexists\ m, a' \cdot m \in M, \kappa(m) = m' \cdot q \xrightarrow{tr|c\mapsto a'}_{m} q')\ \textit{and}\ (q = q'))
\end{aligned}$$

Proof. Each transition $t' \in t'_m$ for any $m' \in M'$ is obtained by applying either step $3a_1$, $3a_2$ or 4b.

If t' was obtained by applying step $3a_1$, then it is a new loop transition added by the translation, i.e., its origin and destination states are the same, and given a *ppDATE* $m \in M$ such that $\kappa(m) = m'$, there not exists a transition associated to t' in m. Therefore, the right side of the disjunction holds.

If t' was obtained by applying step $3a_2$, then, given a *ppDATE* $m \in M$ such that $\kappa(m) = m'$, either there exists one transition on m with the same trigger, same condition, and similar action (but without including the if-expression checking the precondition), or t' is a new loop transition added by the translation. In the first case the left side of the disjunction holds, whereas in the the second case the right side of the disjunction holds.

Finally, if t' was obtained by applying step $3b$, then, given a *ppDATE* $m \in M$ such that $\kappa(m) = m'$, m has exactly the same transition. Therefore, the left-hand side of the disjunction holds in these cases. $\square$

Now, we proceed to prove the lemmas.

Lemma 2. *Given a network of ppDATEs* $pn = (M, V, \nu_0, T_{ppd})$*, its translation* $ppd2DATE(pn) = (M', V, \nu_0, T'_d)$*, a trace* $w \in (systemevent \times \Theta_{Sys})^*$*, and the global configurations* (L, ν) *and* $(\tilde{L}, \nu')$*,*

$$\begin{array}{l}
C_{init}(pn) \overset{w}{\Rightarrow}_M (L, \nu) \text{ and } C_{init}(ppd2DATE(pn)) \overset{w}{\Rightarrow}_{M'} (\tilde{L}, \nu') \\
\text{implies} \\
\quad\quad \forall\ m, q, \rho \cdot (m, q, \rho) \in L, m \in M \cdot \\
\quad\quad\quad\quad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}\ \cdot \kappa(m) = m' \textit{ and } q = q' \\
\quad \text{and} \\
\quad\quad \forall\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \in M' \cdot \\
\quad\quad\quad\quad \exists\ m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L \cdot q = q' \\
\quad \text{and} \\
\quad \forall\ m, q, \rho \cdot (m, q, \rho) \in L, m \notin M \cdot \\
\quad\quad\quad\quad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot q = q' \\
\quad \text{and} \\
\quad \forall\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot \\
\quad\quad\quad\quad \exists\ m, q, \rho \cdot (m, q, \rho), m \notin M \in L \cdot q = q' \\
\quad \text{and} \\
\quad\quad \nu = \nu'
\end{array}$$

Proof. We proceed to prove this lemma by induction on the length of the trace w.

– Base case: $w = \varepsilon$ (empty trace)

$$
\begin{array}{l}
C_{init}(pn) \overset{\varepsilon}{\Rightarrow}_M (L, \nu) \textit{ and } C_{init}(ppd2DATE(pn)) \overset{\varepsilon}{\Rightarrow}_{M'} (\tilde{L}, \nu') \\
\textit{implies} \\
\quad \forall\, m, q, \rho \cdot (m, q, \rho) \in L, m \in M \cdot \\
\quad\quad (\exists\, m', q' \cdot (m', q', \emptyset) \in \tilde{L} \cdot \kappa(m) = m' \textit{and } q = q') \\
\quad \textit{and} \\
\quad \forall\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \in M' \cdot \\
\quad\quad\quad \exists\, m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L \cdot q = q' \\
\quad \textit{and} \\
\quad \forall\, m, q, \rho \cdot (m, q, \rho) \in L, m \notin M \cdot \\
\quad\quad\quad \exists\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot q = q' \\
\quad \textit{and} \\
\quad \forall\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot \\
\quad\quad\quad \exists\, m, q, \rho \cdot (m, q, \rho), m \notin M \in L \cdot q = q' \\
\quad\quad \nu = \nu'
\end{array}
$$

By Def. 17 and Def. 21, we know that

$$
\begin{array}{l}
L_0 = L \textit{ and } \nu_0 = \nu \textit{ and } L'_0 = \tilde{L} \textit{ and } \nu_0 = \nu' \\
\textit{implies} \\
\quad \forall\, m, q, \rho \cdot (m, q, \rho) \in L, m \in M \cdot \\
\quad\quad (\exists\, m', q' \cdot (m', q', \emptyset) \in \tilde{L} \cdot \kappa(m) = m' \textit{and } q = q') \\
\quad \textit{and} \\
\quad \forall\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \in M' \cdot \\
\quad\quad\quad \exists\, m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L \cdot q = q' \\
\quad \forall\, m, q, \rho \cdot (m, q, \rho) \in L, m \notin M \cdot \\
\quad\quad\quad \exists\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot q = q' \\
\quad \textit{and} \\
\quad \forall\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot \\
\quad\quad\quad \exists\, m, q, \rho \cdot (m, q, \rho), m \notin M \in L \cdot q = q' \\
\quad \textit{and} \\
\quad\quad \nu = \nu'
\end{array}
$$

where $L_0 = \{(m, q_{0m}, \emptyset) \mid m \in M\}$, and $L'_0 = \{(m', q_{0m'}, \emptyset) \mid m' \in M'\}$.

Next, by substitution with the antecedents we have to prove

$$\begin{aligned}
&(1)\ \forall\, m, q, \rho \cdot (m, q, \rho) \in L_0, m \in M \cdot \\
&\quad \exists\, m', q' \cdot (m', q', \emptyset) \in L'_0 \cdot \kappa(m) = m' \textit{ and } q = q' \\
&\textit{and} \\
&(2)\ \forall\, m', q' \cdot (m', q', \emptyset) \in L'_0, m' \in M' \cdot \\
&\quad \exists\, m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L_0 \cdot q = q' \\
&\textit{and} \\
&(3)\ \forall\, m, q, \rho \cdot (m, q, \rho) \in L_0, m \notin M \cdot \\
&\qquad \exists\, m', q' \cdot (m', q', \emptyset) \in L'_0, m' \notin M' \cdot q = q' \\
&\textit{and} \\
&(4)\ \forall\, m', q' \cdot (m', q', \emptyset) \in L'_0, m' \notin M' \cdot \\
&\qquad \exists\, m, q, \rho \cdot (m, q, \rho), m \notin M \in L_0 \cdot q = q' \\
&\textit{and} \\
&(5)\ \nu_0 = \nu_0
\end{aligned}$$

As in L'_0 all the *DATE* components of the local configurations correspond to the translation of *ppDATE* in *pn*, both (1) and (2) are trivially fulfilled, and the ranges of both (3) and (4) are never fulfilled, meaning that, as these ranges are empty (i.e., *false*), both expressions are trivially evaluated to *true*. In addition, (5) is trivially fulfilled. Thereby, the base case holds.

– Inductive case: $w = w' : (e, \theta)$

IH: $\forall\, L, \tilde{L}, \nu, \nu' \cdot$

$$\begin{aligned}
&C_{init}(pn) \overset{w'}{\Longrightarrow}_M (L, \nu) \textit{ and } C_{init}(ppd2DATE(pn)) \overset{w'}{\Longrightarrow}_{M'} (\tilde{L}, \nu') \\
&\textit{implies} \\
&\quad \forall\, m, q, \rho \cdot (m, q, \rho) \in L, m \in M \cdot \\
&\qquad \exists\, m', q' \cdot (m', q', \emptyset) \in \tilde{L} \cdot \kappa(m) = m' \textit{ and } q = q' \\
&\quad \textit{and} \\
&\quad \forall\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \in M' \cdot \\
&\qquad \exists\, m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L \cdot q = q' \\
&\quad \textit{and} \\
&\quad \forall\, m, q, \rho \cdot (m, q, \rho) \in L, m \notin M \cdot \\
&\qquad\quad \exists\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot q = q' \\
&\quad \textit{and} \\
&\quad \forall\, m', q' \cdot (m', q', \emptyset) \in L', m' \notin M' \cdot \\
&\qquad\quad \exists\, m, q, \rho \cdot (m, q, \rho), m \notin M \in \tilde{L} \cdot q = q' \\
&\quad \textit{and} \\
&\qquad \nu = \nu'
\end{aligned}$$

Given the previous inductive hypothesis *IH*, we have to prove,

$$
\begin{array}{l}
C_{init}(pn) \xRightarrow{w':(e,\theta)}_{M} (L,\nu) \textit{ and } C_{init}(ppd2DATE(pn)) \xRightarrow{w':(e,\theta)}_{M'} (\tilde{L},\nu') \\
\textit{implies} \\
\quad \forall\, m,q,\rho \cdot (m,q,\rho) \in L, m \in M \cdot \\
\qquad\quad \exists\, m',q' \cdot (m',q',\emptyset) \in \tilde{L} \cdot \kappa(m) = m' \textit{ and } q = q' \\
\quad \textit{and} \\
\quad \forall\, m',q' \cdot (m',q',\emptyset) \in \tilde{L}, m' \in M' \cdot \\
\qquad\qquad \exists\, m,q,\rho \cdot m \in M, \kappa(m) = m', (m,q,\rho) \in L \cdot q = q' \\
\quad \textit{and} \\
\quad \forall\, m,q,\rho \cdot (m,q,\rho) \in L, m \notin M \cdot \\
\qquad\qquad \exists\, m',q' \cdot (m',q',\emptyset) \in \tilde{L}, m' \notin M' \cdot q = q' \\
\quad \textit{and} \\
\quad \forall\, m',q' \cdot (m',q',\emptyset) \in L', m' \notin M' \cdot \\
\qquad\qquad \exists\, m,q,\rho \cdot (m,q,\rho), m \notin M \in \tilde{L} \cdot q = q' \\
\quad \textit{and} \\
\qquad \nu = \nu'
\end{array}
$$

By Def. 21 we have,

$$
\begin{array}{l}
\quad (i)\ \exists\, L'',\nu'' \cdot C_{init}(pn) \overset{w'}{\Rightarrow} (L'',\nu'') \textit{ and } (L'',\nu'') \xRightarrow{(e,\theta)} (L,\nu) \\
\textit{and} \\
\quad (ii)\ \exists\, L'',\nu'' \cdot \\
\qquad\quad C_{init}(ppd2DATE(pn)) \overset{w'}{\Rightarrow} (L'',\nu'') \textit{ and } (L'',\nu'') \xRightarrow{(e,\theta)} (\tilde{L},\nu')
\end{array}
$$

Then, we proceed with the proof by assuming the antecedent of the implication. This assumption allows us to remove the existential quantifiers in the antecedents by introducing the fresh values L'' and ν'' in (i), and the fresh values $\tilde{L}''$ and ν''' in (ii). Therefore, we have

$$
\begin{array}{l}
\quad (i')\ C_{init}(pn) \overset{w'}{\Rightarrow} (L'',\nu'') \textit{ and } (L'',\nu'') \xRightarrow{(e,\theta)} (L,\nu) \\
\textit{and} \\
\quad (ii')\ C_{init}(ppd2DATE(pn)) \overset{w'}{\Rightarrow} (\tilde{L}'',\nu''') \textit{ and } (\tilde{L}'',\nu''') \xRightarrow{(e,\theta)} (\tilde{L},\nu')
\end{array}
$$

Next, by *IH* we know

$(iii)\ \forall\ m, q, \rho \cdot (m, q, \rho) \in L'', m \in M \cdot$
$\quad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}'' \cdot \kappa(m) = m'\ \textit{and}\ q = q'$
and
$(iv)\ \forall\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}'', m' \in M' \cdot$
$\quad \exists\ m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L'' \cdot q = q'$
and
$(v)\ \forall\ m, q, \rho \cdot (m, q, \rho) \in L, m \notin M \cdot$
$\quad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot q = q'$
and
$(vi)\ \forall\ m', q' \cdot (m', q', \emptyset) \in L', m' \notin M' \cdot$
$\quad \exists\ m, q, \rho \cdot (m, q, \rho), m \notin M \in \tilde{L} \cdot q = q'$
and
$(vii)\ \nu'' = \nu'''$

In relation to L, by (i) we know it is obtained from L'' after performing a big step with (e, θ). Thereby, the local configurations on L are either the same as in L'', a modified version of the ones in L'', or new local configurations added to control a *DATE* which is a new instance of a template.

Let us introduce the sets L_{nc}, L_c and L_{new}, to represent the local configurations in each one of the previous categories, respectively. Then, we know that

$$(viii)\ L = L_{nc} \cup L_c \cup L_{new}$$

In addition, by using a similar approach with $\tilde{L}$ and (ii), we introduce the following sets.

$$(ix)\ \tilde{L} = \tilde{L}_{nc} \cup \tilde{L}_c \cup \tilde{L}_{new}$$

Let us come back now to the expression we want to prove.

$(x)\ \forall\ m, q, \rho \cdot (m, q, \rho) \in L, m \in M \cdot$
$\quad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L} \cdot \kappa(m) = m'\ \textit{and}\ q = q'$
and
$(xi)\ \forall\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \in M' \cdot$
$\quad \exists\ m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L \cdot q = q'$
and
$(xii)\ \forall\ m, q, \rho \cdot (m, q, \rho) \in L, m \notin M \cdot$
$\quad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot q = q'$
and
$(xiii)\ \forall\ m', q' \cdot (m', q', \emptyset) \in L', m' \notin M' \cdot$
$\quad \exists\ m, q, \rho \cdot (m, q, \rho), m \notin M \in \tilde{L} \cdot q = q'$
and
$(xiv)\ \nu = \nu'$

By (iii) and (iv), as the values in both L_{nc} and $\tilde{L}_{nc}$ are the same as in L'' and $\tilde{L}''$, respectively, we know that these values fulfil all the previous expressions. Thereby, we can reduce $(viii)$ and (ix) to

$$(viii')\ L = L_c \cup L_{new} \quad (ix')\ \tilde{L} = \tilde{L}_c \cup \tilde{L}_{new}$$

Regarding the newly created local configurations in both L_{new} and $\tilde{L}_{new}$, they do not fulfil the ranges of the universal quantifications in neither (x) nor (xi). In addition, by Prop. 1 and Prop. 2, we know that the only difference in the executed actions in the *ppDATE*s in *pn* and their translation is that the actions in the *DATE*s may include the creation of an instance of template *exit_cond_checker*. Besides, by step 4 in the translation algorithm, we now that both the *ppDATE*s templates and their translations have similar transitions and are initialised in the same state. Thus, (xii) and (xii) are fulfilled for these values, and we can reduce $(viii)$ and (ix) to

$$(viii'')\ L = L_c \quad (ix'')\ \tilde{L} = \tilde{L}_c$$

Therefore, we have to prove,

$$\begin{aligned}
&(x')\ \forall\ m, q, \rho \cdot (m, q, \rho) \in L_c, m \in M \cdot \\
&\qquad\qquad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}_c \cdot \kappa(m) = m' \textit{ and } q = q' \\
&\textit{and} \\
&(xi')\ \forall\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}_c, m' \in M' \cdot \\
&\qquad\quad \exists\ m, q, \rho \cdot m \in M, \kappa(m) = m', (m, q, \rho) \in L_c \cdot q = q' \\
&\textit{and} \\
&(xii')\ \forall\ m, q, \rho \cdot (m, q, \rho) \in L, m \notin M \cdot \\
&\qquad\quad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}_c, m' \notin M' \cdot q = q' \\
&\textit{and} \\
&(xiii')\ \forall\ m', q' \cdot (m', q', \emptyset) \in L_c, m' \notin M' \cdot \\
&\qquad\quad \exists\ m, q, \rho \cdot (m, q, \rho), m \notin M \in \tilde{L}_c \cdot q = q' \\
&\textit{and} \\
&(xiv)\ \nu = \nu'
\end{aligned}$$

By (iii) and Prop. 1 we know that for every enabled transition of a *ppDATE* $m \in M$, there is one enabled transition in $\kappa(m) \in M'$ performing the same change of state and, if any, generating the same action events. Thereby, both *pn* and its translation will shift the local configurations in L_c and $\tilde{L}_c$, respectively, in the same manner, i.e., (x') holds.

In addition, by (iv) and Prop. 2 we know that for every enabled transition in a *DATE* $m' \in M'$, there is either an enabled transition in a *ppDATE* $m \in M$, where $\kappa(m) = m'$, such that this transition performs the same change of state and, if any, generates the same action events, or the transition enabled in m' is a loop transition.

In the first case, both *pn* and its translation will shift the local configurations in L_c and $\tilde{L}_c$, respectively, in the same manner. Thus, (xi') holds.

In the second case, the local configuration obtained after the shift is in the same state as before the shift. Thus, by (iv), this (xi') holds.

Moreover, by *IH*, Prop. 1, Prop. 2 we know that whenever a *ppDATE* in *pn* creates an instance of a template, its translation will create an instance of the translation of such template, and vice versa. Besides, by the step 4 in the translation algorithm, as such instances have similar transitions, they will shift the local configuration associated to them in the same manner. Therefore, both (xii') and $(xiii')$ are fulfilled.

Finally, in relation to (xiv), by Prop. 1 and Prop. 2 we know that only difference in the executed actions in *pn* and its translation is that the actions of the latter may include the creation of an instance of template *exit_cond_checker* (whose actions do not modify *ppDATE* variables valuations). In addition, by step 4 in the translation algorithm we know that both an instance of a *ppDATE* template and a similar instance of the translation of the template will fire similar transitions (with the same actions). Therefore, they perform the same modifications in the valuations ν'' and ν'''. Thus, by (vii), (xiv) holds. □

Lemma 3. *Given a network of* ppDATEs *$pn = (M, V, \nu_0, T_{ppd})$, its translation $ppd2DATE(pn) = (M', V, \nu_0, T'_d)$, a trace $w \in (systemevent \times \Theta_{Sys})^*$, and the global configurations (L, ν) and $(\tilde{L}, \nu')$,*

$$C_{init}(pn) \overset{w}{\Rightarrow}_M (L, \nu) \text{ and } C_{init}(ppd2DATE(pn)) \overset{w}{\Rightarrow}_{M'} (\tilde{L}, \nu') \text{ implies } \psi(L, \tilde{L})$$

where,

$$\begin{aligned}
\psi(L, \tilde{L}) = \;& \forall\, m, q, \rho \cdot (m, q, \rho) \in L \;\cdot \\
& \qquad \forall\, \sigma_{id}^{\uparrow}, \pi', \theta \cdot (\sigma_{id}^{\uparrow}, \pi', \theta) \in \rho \;\cdot \\
& \qquad \exists\, m', q' \cdot (m', q', \emptyset) \in \tilde{L} \cdot inst\,(exit\_cond\_checker, \sigma, \pi') = m' \\
& \quad and \\
& \qquad \forall\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}, m' \notin M' \cdot \\
& \qquad\qquad \exists\, \sigma_{id}^{\uparrow}, \pi' \cdot inst\,(exit\_cond\_checker, \sigma, \pi') = m' \\
& \qquad\qquad\qquad \text{implies } \exists\, m, q, \rho, \theta \cdot (m, q, \rho) \in L \cdot (\sigma_{id}^{\uparrow}, \pi', \theta) \in \rho
\end{aligned}$$

Proof. We proceed to prove this lemma by induction on the length of the trace w.

– Base case: $w = \varepsilon$ (empty trace)

$$C_{init}(pn) \overset{\varepsilon}{\Rightarrow}_M (L, \nu) \textit{ and } C_{init}(ppd2DATE(pn)) \overset{\varepsilon}{\Rightarrow}_{M'} (\tilde{L}, \nu') \textit{ implies } \psi(L, \tilde{L})$$

By Def. 17 and Def. 21 we know that

$$L_0 = L \textit{ and } \nu_0 = \nu \textit{ and } L'_0 = \tilde{L} \textit{ and } \nu_0 = \nu' \textit{ implies } \psi(L, \tilde{L})$$

where $L_0 = \{(m, q_{0m}, \emptyset) \mid m \in M\}$, and $L'_0 = \{(m', q_{0m'}, \emptyset) \mid m' \in M'\}$.

Next, by substitution with the antecedents,

$$L_0 = L \textit{ and } \nu_0 = \nu \textit{ and } L'_0 = \tilde{L} \textit{ and } \nu_0 = \nu' \textit{ implies } \psi(L_0, L'_0)$$

Thus, by the definition of ψ we have to prove that,

$$\begin{array}{l}
\forall\ m, q, \rho \cdot (m, q, \rho) \in \{(m, q_{0m}, \emptyset) \mid m \in M\}\ \cdot \\
\quad \forall\ \sigma^{\uparrow}_{id}, \pi', \theta \cdot (\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho\ \cdot \\
\qquad \exists\ m', q' \cdot (m', q', \emptyset) \in \{(m', q_{0m'}, \emptyset) \mid m' \in M'\} \cdot \\
\qquad\quad \textit{inst}\,(\textit{exit\_cond\_checker}, \sigma, \pi') = m' \\
\textit{and} \\
\forall\ m', q' \cdot (m', q', \emptyset) \in \{(m', q_{0m'}, \emptyset) \mid m' \in M'\}, m' \notin M' \cdot \\
\quad \exists\ \sigma^{\uparrow}_{id}, \pi' \cdot \textit{inst}\,(\textit{exit\_cond\_checker}, \sigma, \pi') = m' \\
\qquad\qquad \textit{implies}\ \exists\ m, q, \rho, \theta \cdot (m, q, \rho) \in \{(m, q_{0m}, \emptyset) \mid m \in M\} \cdot \\
\qquad\qquad\qquad (\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho
\end{array}$$

First, let us analyse the expression,

$$\begin{array}{l}
\forall\ m, q, \rho \cdot (m, q, \rho) \in \{(m, q_{0m}, \emptyset) \mid m \in M\}\ \cdot \\
\quad \forall\ \sigma^{\uparrow}_{id}, \pi', \theta \cdot (\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho\ \cdot \\
\qquad \exists\ m', q' \cdot (m', q', \emptyset) \in \{(m', q_{0m'}, \emptyset) \mid m' \in M'\} \cdot \\
\qquad\quad \textit{inst}\,(\textit{exit\_cond\_checker}, \sigma, \pi') = m'
\end{array}$$

As ρ is always the empty set, the condition $(\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho$ will always evaluate to *false*. Therefore,

$$\begin{array}{l}
\forall\ m, q, \rho \cdot (m, q, \rho) \in \{(m, q_{0m}, \emptyset) \mid m \in M\}\ \cdot \\
\quad \forall\ \sigma^{\uparrow}_{id}, \pi', \theta \cdot false \cdot \\
\qquad \exists\ m', q' \cdot (m', q', \emptyset) \in \{(m', q_{0m'}, \emptyset) \mid m' \in M'\} \cdot \\
\qquad\quad \textit{inst}\,(\textit{exit\_cond\_checker}, \sigma, \pi') = m'
\end{array}$$

Then, as the range of the inner universal quantification is empty (i.e., *false*), it is trivially evaluated to *true*.

$$\forall\ m, q, \rho \cdot (m, q, \rho) \in \{(m, q_{0m}, \emptyset) \mid m \in M\} \cdot \textit{true}$$

Finally, as the body of the previous universal quantification is simply the value *true* and its range is not empty, the whole expression is trivially evaluated to *true*.

Now, let us analyse the expression,

$$\begin{aligned}&\forall\ m', q' \cdot (m', q', \emptyset) \in \{(m', q_{0m'}, \emptyset) \mid m' \in M'\}, m' \notin M' \cdot\\ &\quad \exists\ \sigma^{\uparrow}_{id}, \pi' \cdot \mathit{inst}\,(\mathit{exit\_cond\_checker}, \sigma, \pi') = m'\\ &\qquad\qquad \mathit{implies}\ \exists\ m, q, \rho, \theta \cdot (m, q, \rho) \in \{(m, q_{0m}, \emptyset) \mid m \in M\} \cdot\\ &\qquad\qquad\qquad (\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho\end{aligned}$$

As in the initial configuration of the translation of *pn* there are no instances of *DATE* templates, the range of the universal quantification is always evaluated to *false*. Therefore,

$$\begin{aligned}&\forall\ m', q' \cdot false \cdot\\ &\quad \exists\ \sigma^{\uparrow}_{id}, \pi' \cdot \mathit{inst}\,(\mathit{exit\_cond\_checker}, \sigma, \pi') = m'\\ &\qquad\qquad \mathit{implies}\ \exists\ m, q, \rho, \theta \cdot (m, q, \rho) \in \{(m, q_{0m}, \emptyset) \mid m \in M\} \cdot\\ &\qquad\qquad\qquad (\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho\end{aligned}$$

Thus, as the range of the universal quantification is empty (i.e., *false*), the whole expression is trivially evaluated to *true*. Thereby, the base case holds.

– Inductive case: $w = w' : (e, \theta)$

IH: $\forall\ L, \tilde{L}, \nu, \nu'\ \cdot$

$$\begin{aligned}&C_{\mathit{init}}(\mathit{pn}) \overset{w'}{\Longrightarrow}_{M} (L, \nu)\ \mathit{and}\ C_{\mathit{init}}(\mathit{ppd2DATE}(\mathit{pn})) \overset{w'}{\Longrightarrow}_{M'} (\tilde{L}, \nu')\\ &\mathit{implies}\ \psi(L, \tilde{L})\end{aligned}$$

Given the previous inductive hypothesis *IH*, we have to prove,

$$\begin{aligned}&C_{\mathit{init}}(\mathit{pn}) \xRightarrow{w':(e,\theta)}_{M} (L, \nu)\ \mathit{and}\ C_{\mathit{init}}(\mathit{ppd2DATE}(\mathit{pn})) \xRightarrow{w':(e,\theta)}_{M'} (\tilde{L}, \nu')\\ &\quad \mathit{implies}\ \psi(L, \tilde{L})\end{aligned}$$

By Def. 21 we have,

$$\begin{aligned}&\quad (i)\ \exists\ L'', \nu'' \cdot C_{\mathit{init}}(\mathit{pn}) \overset{w'}{\Longrightarrow} (L'', \nu'')\ \mathit{and}\ (L'', \nu'') \xRightarrow{(e,\theta)} (L, \nu)\\ &\mathit{and}\\ &\quad (ii)\ \exists\ L'', \nu'' \cdot C_{\mathit{init}}(\mathit{ppd2DATE}(\mathit{pn})) \overset{w'}{\Longrightarrow} (L'', \nu'')\\ &\qquad \mathit{and}\ (L'', \nu'') \xRightarrow{(e,\theta)} (\tilde{L}, \nu')\ \mathit{implies}\ \psi(L, \tilde{L})\end{aligned}$$

Then, we proceed with the proof by assuming the antecedent of the implication. This assumption allows us to remove the existential quantifiers in the antecedents by introducing the fresh values L'' and ν'' in (i), and the fresh values $\tilde{L}''$ and ν''' in (ii). Therefore, we have

$$\begin{aligned}&\quad (i')\ C_{\mathit{init}}(\mathit{pn}) \overset{w'}{\Longrightarrow} (L'', \nu'')\ \mathit{and}\ (L'', \nu'') \xRightarrow{(e,\theta)} (L, \nu)\\ &\mathit{and}\\ &\quad (ii')\ C_{\mathit{init}}(\mathit{ppd2DATE}(\mathit{pn})) \overset{w'}{\Longrightarrow} (\tilde{L}'', \nu''')\ \mathit{and}\ (\tilde{L}'', \nu''') \xRightarrow{(e,\theta)} (\tilde{L}, \nu')\end{aligned}$$

Next, by *IH* we know that $\psi(L'', \tilde{L}'')$. Thus, we have

$$(iii)\ \psi(L'', \tilde{L}'')$$

In relation to L, by (i') we know it is obtained from L'' after performing a big step with (e, θ). Thereby, the local configurations on L are either the same as in L'', a modified version of the ones in L'', or new local configurations added to control a *DATE* which is a new instance of a template.

Let us introduce the sets L_{nc}, L_c and L_{new}, to represent the local configurations in each one of the previous categories, respectively. Then, we know that

$$(iv)\ L = L_{nc} \cup L_c \cup L_{new}$$

In addition, by using a similar approach with $\tilde{L}$ and (ii'), we introduce the following sets.

$$(v)\ \tilde{L} = \tilde{L}_{nc} \cup \tilde{L}_c \cup \tilde{L}_{new}$$

As in the translation the set $\tilde{L}_{new}$ contains both the instances of ordinary templates and the instances of the templates about Hoare triples, we split $\tilde{L}_{new}$ into the sets $\tilde{L}'_{new}$ and $\tilde{L}_h$, to represent each one of the previous categories, respectively. Thus,

$$(v')\ \tilde{L} = \tilde{L}_{nc} \cup \tilde{L}_c \cup \tilde{L}'_{new} \cup \tilde{L}_h$$

Now, let us come back to the expression $\psi(L, \tilde{L})$. By (iv) and (v'), we replace it by

$$\psi(L_{nc} \cup L_c \cup L_{new}, \tilde{L}_{nc} \cup \tilde{L}_c \cup \tilde{L}'_{new} \cup \tilde{L}_h)$$

By (iii), as the values in both L_{nc} and $\tilde{L}_{nc}$ are the same as in L'' and $\tilde{L}''$, respectively, we know that the former fulfil ψ. Thereby, we can reduce the previous expression to

$$\psi(L_c \cup L_{new}, \tilde{L}_c \cup \tilde{L}'_{new} \cup \tilde{L}_h)$$

In addition, newly created local configurations in both L_{new} and $\tilde{L}'_{new}$ do not fulfil the ranges of the quantified expressions in ψ. Then, we can discard them.

$$\psi(L_c, \tilde{L}_c \cup \tilde{L}_h)$$

Next, by the definition of ψ, we have

$$\begin{array}{l}
(vi)\ \forall\, m, q, \rho \cdot (m, q, \rho) \in L_c \cdot \\
\qquad \forall\, \sigma_{id}^{\uparrow}, \pi', \theta \cdot (\sigma_{id}^{\uparrow}, \pi', \theta) \in \rho \cdot \\
\qquad \exists\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}_c \cup \tilde{L}_h \cdot inst\,(exit\_cond\_checker, \sigma, \pi') = m' \\
and \\
(vii)\ \forall\, m', q' \cdot (m', q', \emptyset) \in \tilde{L}_c \cup \tilde{L}_h, m' \notin M' \cdot \\
\qquad \exists\, \sigma_{id}^{\uparrow}, \pi' \cdot inst\,(exit\_cond\_checker, \sigma, \pi') = m' \\
\qquad\qquad implies\ \exists\, m, q, \rho, \theta \cdot (m, q, \rho) \in L_c \cdot (\sigma_{id}^{\uparrow}, \pi', \theta) \in \rho
\end{array}$$

In relation to the configurations in $\tilde{L}_c$, as they were obtained from configurations in $\tilde{L}''$, by (*iii*) we know they fulfil (*vii*) (same *DATE* component). Thereby, we only need to prove that

$$\begin{aligned}
&(vi)\ \forall\ m, q, \rho \cdot (m, q, \rho) \in L_c \cdot \\
&\qquad \forall\ \sigma_{id}^{\uparrow}, \pi', \theta \cdot (\sigma_{id}^{\uparrow}, \pi', \theta) \in \rho \cdot \\
&\qquad \exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}_c \cup \tilde{L}_h \cdot \mathit{inst}\,(\mathit{exit\_cond\_checker}, \sigma, \pi') = m' \\
&\mathit{and} \\
&(vii')\ \forall\ m', q' \cdot (m', q', \emptyset) \in \tilde{L}_h, m' \notin M' \cdot \\
&\qquad \exists\ \sigma_{id}^{\uparrow}, \pi' \cdot \mathit{inst}\,(\mathit{exit\_cond\_checker}, \sigma, \pi') = m' \\
&\qquad\qquad \mathit{implies}\ \exists\ m, q, \rho, \theta \cdot (m, q, \rho) \in L_c \cdot (\sigma_{id}^{\uparrow}, \pi', \theta) \in \rho
\end{aligned}$$

Now, let us focus on (*vi*). If event e is either an exit event, or an entry event which does not require to verify any Hoare triple, then it does not introduce any new values in ρ components of the local configurations in L_c. Thus, by (*iii*), (*vi*) is fulfilled in both cases.

If event e is an entry event which requires the check of Hoare triples, then by Lemma 2 and Prop. 1, we know that for every enabled transition which requires the verification of a Hoare triple in *pn*, a similar transition will be fired in its translation whose action will create a *DATE* in charge of controlling such Hoare triple. Thus, for every new entry in a ρ component in L_c, a new local configuration is added in $\tilde{L}_h$. Thereby, (*vi*) holds.

Regarding (*vii'*), if event e is either an exit event, or an entry event which does not require to verify any Hoare triple, then $\tilde{L}_h = \emptyset$. Thus, as the range of universal quantification is empty, (*vii'*) is trivially fulfilled in both cases.

If event e is an entry event which requires the check of Hoare triples, then by the rules entry_1 and entry_3 in the relation *small step local*, we know that a new tuple is going to be added to the ρ component of the local configuration in L_c which are associated to the *ppDATE*s whose current state possess a Hoare triple that has to be verified. In addition, a local configuration is going to be included in $\tilde{L}_h$ for the *DATE* instantiated to control the corresponding Hoare triple. Thereby, (*vii'*) holds. □

B Proof of Soundness

Theorem 1. *Given a ppDATE network* $pn = (M, V, \nu_0, T_{ppd})$*, and its translation* $ppd2DATE(pn) = (M', V, \nu_0, T'_d)$*,*

$$\mathcal{VT}(pn) = \mathcal{VT}(ppd2DATE(pn))$$

Proof. To prove this theorem we will show that,

$$\forall w \cdot w \in (\mathit{systemevent} \times \Theta_{Sys})^* \cdot w \in \mathcal{VT}(pn)\ \mathit{iff}\ w \in \mathcal{VT}(ppd2DATE(pn))$$

In the following, we abbreviate *ppd2DATE*(*pn*) by *dn*.

– *$w \in \mathcal{VT}(pn)$ implies $w \in \mathcal{VT}(dn)$*

As $w \in \mathcal{VT}(pn)$, by Def. 22 we know that it has a prefix w' such that either,

(i) $C_{init}(pn) \overset{w'}{\Longrightarrow}_M (L', \nu')$ and $\exists\ (m, q, \rho) \cdot (m, q, \rho) \in L' \cdot q \in B_m$, or
(ii) $w' = w_1 +\!\!+ \langle(\sigma^{\uparrow}_{id}, \theta')\rangle$, $C_{init}(pn) \overset{w_1}{\Longrightarrow} (L', \nu')$ and $\exists\ m, q, \rho, \pi', \theta \cdot ((m, q, \rho) \in L' \mathit{and}\ (\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho) \cdot \theta, \theta' \not\models \pi'$.

In relation to (i), let us assume that exists $(\tilde{L}, \nu)$ such that $C_{init}(dn) \overset{w'}{\Longrightarrow}_{M'} (\tilde{L}, \nu)$. Then, by Lemma 2 we know that for every local configuration in L', there is a local configuration in $\tilde{L}$ such that its state component is the same. Therefore, as in L' there is a local configuration in a bad state, there is a local configuration in $\tilde{L}$ in a bad state, i.e. w' is a counter-example of dn. Thereby, $w \in \mathcal{VT}(dn)$.

Regarding (ii), it corresponds to the case where (at least) one Hoare triple is not fulfilled when event $\sigma^{\uparrow}_{id}$ occurs. Here, by Lemma 3 we have

$$\psi(L', \tilde{L})$$

Therefore, by (ii) and $\psi(L', \tilde{L})$ we know that

$$\exists\ m', q' \cdot (m', q', \emptyset) \in \tilde{L} \cdot \mathit{inst}(\mathit{exit\_cond\_checker}, \sigma, \texttt{part\_eval}(\pi')) = m'$$

Let us assume that the local configuration $(m', q', \emptyset)$ is the one satisfying the previous existential quantification. In addition, let us assume $(\tilde{L}, \nu)$ to be given by $C_{init}(dn) \overset{w_1}{\Longrightarrow}_{M'} (\tilde{L}, \nu)$. Then, once $\sigma^{\uparrow}_{id}$ occurs, as by (ii) we know that the π' is not fulfilled, m' will shift to a bad state. Thereby, w' is a counter-example of dn, i.e. $w \in \mathcal{VT}(dn)$.

– *$w \in \mathcal{VT}(dn)$ implies $w \in \mathcal{VT}(pn)$.*

As $w \in \mathcal{VT}(dn)$, by Def. 22 and the fact that every *DATE* in dn has no Hoare triples associated to its states, we know that it has a prefix w' such that,

$$C_{init}(dn) \overset{w'}{\Longrightarrow}_{M'} (\tilde{L}, \nu)\ \mathit{and}\ \exists\ (m, q, \rho) \cdot (m, q, \rho) \in \tilde{L} \cdot q \in B_m$$

Now let us assume that exists (L', ν') such that $C_{init}(pn) \overset{w'}{\Longrightarrow}_M (L', \nu')$. In addition, let us assume that the bad state in $\tilde{L}$ belongs to a local configuration associated to a *DATE* m', which is an instance of the template *exit_cond_checker*, i.e., m' was created to control a Hoare triple.

Let us represent this Hoare triple as $\{\pi\}\,\sigma\,\{\pi'\}$. Then, by Lemma 3 we know that,

$$(1)\ \exists\ m, q, \rho, \theta \cdot (m, q, \rho) \in L' \cdot (\sigma^{\uparrow}_{id}, \pi', \theta) \in \rho$$

We will assume that the *ppDATE* m and the valuation θ are the ones fulfilling (1). Note that the index *id* is introduced by Lemma 3. Next, as m' is in a bad state we know that whenever $\sigma^{\uparrow}_{id}$ occurs, π' is not fulfilled. Thus, let us assume that the selected prefix is of the form $w' = w_1 +\!\!+ \langle(\sigma^{\uparrow}_{id}, \theta')\rangle$. Thereby, by Def. 22, w' is a counter-example of pn, i.e. $w \in \mathcal{VT}(pn)$.

On the other hand, if the bad state in $\tilde{L}$ does not belongs to a local configuration associated to a *DATE* m' which is an instance of the template *exit_cond_checker*, then by Lemma 2 we know that there is a local configuration in L' such that its state component is the same as the bad state in $\tilde{L}$. Therefore, w' is a counter-example of pn, i.e. $w \in \mathcal{VT}(pn)$. $\square$

CHAPTER

THREE

STARVOORS USER MANUAL

Jesús Mauricio Chimento

1 Introduction

Day by day the use of formal verification techniques to verify the correctness of programs is increasing. In general, verification tools use either static verification techniques (i.e., the verification is performed prior to program execution), or dynamic verification techniques (i.e., the verification is performed during program execution), in order to verify whether a program fulfills certain properties.

Nowadays, a new trend focused on the combination of static and dynamic verification techniques is starting to emerge. StaRVOOrS (STAtic and Runtime Verification of Object-ORiented Software) is a tool which aims at both the specification and verification of properties by combining the use of *Static Verification* and *Runtime Verification.* On the whole, StaRVOOrS is fed with a Java program and a *ppDATE* specification [5] describing properties which the program under scrutiny must fulfill, and it automatically generates a runtime monitor which will verify the specified properties (at runtime) whenever the provided program is executed.

This document is the user manual of StaRVOOrS. Its structure is as follows. Section 2 provides an intuitive description of the *ppDATE* specification language used by this tool. Section 3 gives a high level explanation about how this tool works. Section 4 shows how to write a *ppDATE* specification in the input language of the tool. Finally, section 5 provides a complete example on how to run this tool.

2 *ppDATE* Specification Language

Here, we briefly introduce the *ppDATE* specification language. However, its complete description can be found in [5].

ppDATE is an automaton-based formalism which, basically, consists of a transition system whose states may include Hoare triples describing properties about the methods of the system under scrutiny.

Transitions in a *ppDATE* are labelled by a trigger (*tr*), a condition (c) and an action (a). Together, the label is written $tr \mid c \mapsto a$. A transition is *enabled* to be taken whenever its trigger is active and the condition guarding it holds. In addition, if a transition is taken, we say that it is *fired.* Whenever a transition is fired, its action is executed.

Regarding the triggers, they are activated by the occurrence of either a visible *system event* such as entering or exiting a method, or an *action event* generated by certain actions labelling other transitions. We use the notation $\texttt{foo}^{\downarrow}$, $\texttt{foo}^{\uparrow}$, $e?$, to represent the trigger which is activated whenever the method `foo` is entered, the trigger which is activated whenever the method `foo` is exited, and the trigger which is activated whenever the action event e occurs, respectively.

Regarding the conditions, they are expressions written using JML boolean expression syntax [9]. Conditions may depend on the values of

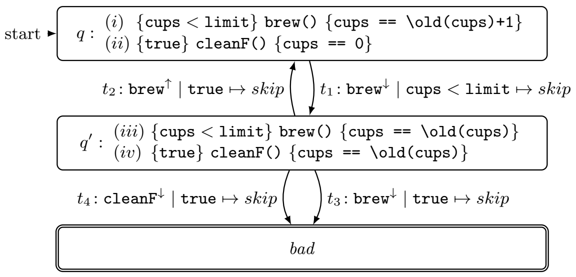

start
q : (i) {cups < limit} brew() {cups == \old(cups)+1}
(ii) {true} cleanF() {cups == 0}
t2 : brew↑ | true ↦ skip
t1 : brew↓ | cups < limit ↦ skip
q′ : (iii) {cups < limit} brew() {cups == \old(cups)}
(iv) {true} cleanF() {cups == \old(cups)}
t4 : cleanF↓ | true ↦ skip
t3 : brew↓ | true ↦ skip
bad

Fig. 1. A *ppDATE* controlling the brew of coffee

system variables (i.e., of the system under scrutiny) and the values of *ppDATE variables* (i.e., variables which belong to the *ppDATE*). The latter can be modified via actions in the transitions.

Regarding the actions, they consist on any number of the following: (i) assignments of the form $v = exp$, where v is a *ppDATE* variable and *exp* is an expression that may depend on system variables and *ppDATE* variables; (ii) an action *!* such that *e!* represents the generation of the action event *e*; (iii) an action *create*, used to generate instances of a *ppDATE* template (see Sec. 4.3); (iv) IF-THEN conditional expressions whose branching condition depends on the valuations of system variables and *ppDATE* variables; (v) an action *log* such that *log(string)* adds *string* into the log file generated by the monitor; (vi) and (Java) programs. All the actions should end in a semicolon.

In relation to the Hoare triples on the states of a *ppDATE*, intuitively, if a Hoare triple $\{\pi\}\, foo()\, \{\pi'\}$ is included in some state q, this property ensures that: if method `foo` is entered while the monitor is in state q, and pre-condition π holds, then upon reaching the corresponding exit from `foo`, post-condition π' should hold. Both pre-/post-conditions in the Hoare triples are expressed using *JML Boolean Expressions* syntax (see Sec. 4.5 for details about this syntax).

Now, let us introduce an example in order to give a better intuition on how a *ppDATE* is described.

ppDATE Specification Example

Let us consider a *coffee machine system* where, after a certain amount of coffee cups are brewed, its filters have to be cleaned. If the limit of coffee cups is reached, the machine should not be able to brew any more coffee. In addition, while the coffee machine is active (a coffee cup is being brewed), it is not possible to start brewing another coffee, or to clean the filters.

Fig. 1 illustrates a *ppDATE* describing this part of the system. In other words, whenever the coffee machine is not active, i.e., the machine is not brewing a cup of coffee, and the method `brew` starts the coffee brewing process, then it is not possible either to execute this method again, or to execute the method `cleanF` (which initialises the task of cleaning the filter), until the initialised brewing process finishes.

The previous property can be interpreted as follows: initially being in state q, the state which represents that the coffee machine is not active, whenever method `brew` is invoked and it is possible to brew a cup of coffee (i.e., the limit of coffee cups was not reached yet), then transition t_1 shifts the *ppDATE* from state q to state *q'*. While in *q'*, the state which represents that the coffee machine is active, if either method `brew` or method `cleanF` are invoked, then transitions t_3 or transition t_4 shift the *ppDATE* to state *bad*, respectively. This indicates that the property was violated. On the contrary, if method `brew` terminates its execution, then transition t_2 shifts the *ppDATE* from state *q'* to state q. Note that the names used on the transitions, e.g. t_1, t_2, etc, are not part of the specification language. They are included to simplify the description of how the *ppDATE* works.

In addition to this, the Hoare triples in state q ensure the properties: (i) if the amount of brewed coffee cups has not reached its limit yet, then a coffee cup is brewed; (ii) cleaning the filters sets the amount of brewed coffee cups to 0. Property (i) has to be verified if, while the *ppDATE* is on state q, the method `brew` is executed and its precondition holds; and property (ii) has to be verified if, while the *ppDATE* is on state q, the method `cleanF` is executed and its precondition holds. Regarding state *q'*, the Hoare triples in this state ensure the properties: (iii) no coffee cups are brewed; (iv) filters are not cleaned. Property (iii) and (iv) are verified if either method `brew` and method `cleanF` are executed, and their preconditions hold, respectively. Here, remember that this state represents that the coffee machine is active. Thus, if it occurs that either the method `brew` or the method `cleanF` are executed while the *ppDATE* is on this state, then, as this would move the *ppDATE* to state `bad`, one would expect the value of the variable `cup` to remain unchanged. This is precisely what is verified when either property (iii) or (iv) are analysed.

Note that none of the Hoare triples makes reference to the state of the coffee machine, i.e., there is no information about whether the machine is active or not. This is due to fact that the state of the machine is implicitly defined by the states of the *ppDATE*. If the *ppDATE* is in state q, the coffee machine is not active. However, if it is in state *q'*, then the machine is active. Therefore, the Hoare triples are *context dependent*. This is the reason why, we can describe properties with the same precondition, but with different postconditions depending on the state of the *ppDATE* in which they are placed.

3 High-level Description of StaRVOOrS

StaRVOOrS takes three arguments: (i) the path to the main folder of the Java files to be verified; (ii) a description (as input language) of the *ppDATE* specification for the provided program; and (iii) the path of the output folder (the generated files are stored in this folder). Then, it automatically generates (1) a runtime monitor; (2) an instrumented version of the Java files in (i); (3) a report summarising the results obtained by statically verifying the Hoare triples described in (ii); (4) and a refined version of (ii), when possible.

To generate such output, StaRVOOrS combines the use of the deductive source code verifier KeY [4] with the runtime monitoring tool Larva [7]. KeY is a deductive verification system for data-centric *functional correctness* properties of Java programs, which generates, from JML [9] and Java, proof obligations in *Dynamic Logic* (a modal logic for reasoning about programs) [8], and attempts to prove them by using a *sequent calculus* which follows the *symbolic execution* paradigm. Larva is an automata-based Runtime Verification tool for Java programs which automatically generates a runtime monitor from a property using the automaton-based specification language *DATE* [6]. Larva transforms such specification into monitoring code together with AspectJ code to link the system under scrutiny with the generated monitor.

In a nutshell, StaRVOOrS output is generated by following the steps enumerated below.

(a) The Hoare triples described in (ii) are translated into JML contracts, which are textually added to to the Java files in (i) as annotations of the respective methods;

(b) KeY attempts to (statically) verify all the JML contracts automatically. The result obtained for each contract is either a complete proof, or a partial proof where some parts of the contract are proved and others are not, or that KeY cannot prove any of the parts the contract. These results are stored in a XML file. In addition, a report summarising the content of this file, i.e., (3), is generated. Here, note that our tool does not support user interaction with KeY. It uses this prover in fully automatic mode;

(c) The *ppDATE* specification is refined based on the XML file, i.e., (4). Fully verified Hoare triples are removed from the specification, but those Hoare triples which are not fully verified, are left in the specification to be verified at runtime. However, the original preconditions of the remaining Hoare triples may be strengthen with the (path) conditions resulting from partial proofs, thus covering at runtime only executions that are not closed in the static verification step;

(d) The refined *ppDATE* specification is encoded into a *DATE* specification. In particular, the *DATE* specification language does not support pre/post-conditions which thus have to be translated to use notions native to this specification language. This also requires a number of changes to the system (through code instrumentation), in order to be able to distinguish different executions of the same code unit, and to evaluate the Hoare triples in the states of the refinded *ppDATE* at runtime. i.e, (2). Regarding the former, method declarations get a new argument which is used as a counter for invocations of this method. Regarding the latter, not every condition in a pre/postcondition of a Hoare triple can be directly written as a Java Boolean Expression, e.g., quantified expressions. Thus, methods which operationalise the evaluation of those conditions are added to the Java files in (i);

(e) The Larva compiler generates a runtime monitor using aspect-oriented programming techniques, i.e., (1).

Once deployed, the runtime monitor and the instrumented version of the Java files are executed together, thus effectively running the monitor in parallel with the program. The runtime monitor identifies violations at runtime, reporting error traces to be analysed.

4 Composing a *ppDATE* Specification in the Input Language of StaRVOOrS

In this section we explain in detail how to write a *ppDATE* specification using the input language of StaRVOOrS. The files written in such language have extention *.ppd*, and their content may consist on 6 sections which are ordered as follows: `IMPORTS`, `GLOBAL`, `TEMPLATES`, `CINVARIANTS`, `HTRIPLES` and `METHODS`. Below, we describe the content of each one of these sections, show their syntax, and provide examples illustrating how to write them.

4.1 IMPORTS

Section `IMPORTS` lists the packages included in the sytem under scrutiny which are related to the properties to be verified (both the Hoare triples and the automata). Its syntax is described as follows:

```
IMPORTS { import package ; }
```

Each package listed in this section follows the usual Java syntax for imports. For instance,

```
IMPORTS {
  import main.Foo ;
  import other.sub.Goo ;
  import other.Hoo ;
}
```

4.2 GLOBAL

Section GLOBAL contains the description of the *ppDATE* specification. Its syntax, which is described below, is written as follows:

```
GLOBAL {
 VARIABLES { -- definition of the variables -- }

 ACTEVENTS  { -- definition of the action events -- }

 TRIGGERS { -- definition of the triggers -- }

 PROPERTY property_name1 {
  STATES { -- definition of the states of the ppDATE -- }
  TRANSITIONS { -- definition of the transitions of the ppDATE -- }
 }

 PROPERTY property_name2 {
   -- definition of states and transitions --
 }
 ...
}
```

Note that one may describe more than one PROPERTY. This would be the case when one is describing several *ppDATE*s in one single specification file, i.e., each property represents a *ppDATE*.

VARIABLES Subsection VARIABLES allows to include as part of the specification the declaration of variables. These variables, which are referred to as *ppDATE* variables, may be freely used in the transitions of a *ppDATE*, both in their conditions and actions. For instance, one may use an integer variable as a counter to keep track of how many times a method is executed. Below, we illustrate how variables may be defined within this subsection.

```
VARIABLES {
  type var ;
  type var = initial_value ;
}
```

Such syntax follows the usual Java syntax for the declaration of variables. For instance,

```
VARIABLES {
  String s;
  int i = 0;
}
```

Note that whenever a variable is not initialised when it is defined, its initialisation has to be performed by the execution of an action. Otherwise, there is going to be an exception at runtime whenever the monitor attempts to manipulate such variable.

ACTEVENTS Subsection `ACTEVENTS` includes the declaration of the different *action events* which may be generated by using the action *!*. Here, it is only necessary to list the names of these events, as illustrated in the example below for the action events `e1`, `e2`, and `e3`.

```
ACTEVENTS {
  e1 ; e2 ; e3 ;
}
```

TRIGGERS Subsection `TRIGGERS` includes the declaration of the different triggers which may be used in the transitions of a *ppDATE*.

Triggers Associated to System Events The triggers which are activated by the occurrence of a visible *system event*, i.e., entering or exiting a method, have the following signature:

```
   name(args) = {Class obj.method(args')sysevent }
```

Here, `name` is a label which works as an identifier for the trigger; `method` is the name of the method generating the system event which activates the trigger; `obj` is the target object (instance of the class `Class`) on which `method` is being called[1]; and `sysevent` represents whether the trigger is activated by a system event produced by entering or exiting a method, represented with the notation `entry` or `exit`, respectivily. In addition, each trigger may have a number of arguments `args` which act as binds for `args'` (i.e., `args'` are the arguments in `args`, but without their types). This allows the access at runtime to the arguments which are being provided to the method, and to the value returned by a method (see examples below). Note that one may use `obj` to access at runtime the target object as well.

Below, by considering the Java classes depicted in Fig. 2, we give several examples illustrating the different manners in which this kind of triggers might be defined.

```
TRIGGERS {
  foo1()                  = {Foo f.foo()entry}
  foo2()                  = {*.foo()entry}
  goo1(int x, boolean b) = {Goo g.goo(x,b)entry}
  goo2(int x)             = {Goo g.goo(x,*)entry}
  foo3()                  = {*.foo()exit()}
```

[1] We assume that `obj` is universally quantified.

```
public class Foo {          public class Goo {
   public void foo();          public void goo(int x,boolean b);
}                              public int hoo(int x, int y, int z);
                            }
```

Fig. 2. Example of Java classes.

```
  goo4(int x,boolean b)  = {Goo g.goo(x,b)exit()}
  hoo(int x, int ret)    = {Goo g.hoo(x,*)exit(int ret)}
}
```

On these definitions, whenever either the target object of the method or any of the arguments of the method are not necessary for the definition of a trigger, e.g., the definition of a trigger which is activated by the execution of a method `foo` where several classes have an implementation for this method, they can simply be omitted by replacing them with the symbol '$*$', which is used as a placeholder. Triggers `foo2`, `goo2`, `foo3`, and `hoo` are examples illustrating these situations. In addition, in the definition of a trigger which is activated by a system event produced by exiting a method, it is possible to refer (and later to access) to the value (or object) returned by the method, by including in the arguments of the trigger an argument with an appropiate type to represent such value, and then including this argument in the notation `exit`, as it is illustrated in the definition of trigger `hoo`.

PROPERTY The core of this section is the subsection `PROPERTY`. It consists of the actual description of a *ppDATE*. This subsection is divided in two parts: `STATES` and `TRANSITIONS`. Note that there should be defined at least one property here.

STATES `STATES` lists all of the states in a *ppDATE*. There are four kind of states: starting states, accepting states, bad states, and normal states. `STARTING` list the intial state of the *ppDATE*. There should be only one starting state listed. The accepting states, which are listed in `ACCEPTING`, represent the states in which it is desirable for the monitor to be in whenever the program under scrutiny terminates its execution. The bad states, which are listed in `BAD`, represent states which a monitor reaches whenever a property which is described with the transitions of the *ppDATE* is violated at runtime. Finally, normal states, which are listed in `NORMAL`, are neither accepting nor bad states, but simply possible states where a monitor may be in during the execution of a program.

In relation to the list of states, each entry consists on the name of a state, and a list of the names of the Hoare triples which have to be verified in that state (this is properly explained in Sec. 4.5). Entries in a list of states terminate in a semicolon.

Below you can see an example of a `STATES` subsection.

```
STATES {
 STARTING { q0 ; }
 ACCEPTING { q4 (h1,h4) ; q5 ; q6 (h2); }
 BAD { bad ; }
 NORMAL { q1 ; q2 (h2,h3) ; q3; }
}
```

Note that the order previously illustrated in the syntax (starting, accepting, bad, normal) should be preserved. In addition, it is mandatory to always include a starting state on a *ppDATE*. Otherwise, the monitor will not know from which state it should start.

TRANSITIONS Section `TRANSITIONS` contains the description of all the transitions in the *ppDATE*. For each *ppDATE* transition going from state `q` to state `q'` with trigger `tr`, condition `c`, and action `a`, this section includes a line of the form

```
q -> q' [tr \ c \ a]
```

However, the conditions and the actions are optional. Below we list all the syntactically valid expressions which may be used within this section.

(1) `q -> q' [tr\c\a]`
(2) `q -> q' [tr\\]`
(3) `q -> q' [tr\]`
(4) `q -> q' [tr]`
(5) `q -> q' [tr\c\]`
(6) `q -> q' [tr\c ]`
(7) `q -> q' [tr\\a]`

Note that the expressions (2), (3), and (4) are equivalent. Similarly, expressions (5) and (6) are equivalent as well. In addition, when using a trigger in a transition it is not necessary to write their arguments in the trigger component of the transition. This does not affect the possibility of using such arguments in both the conditions, and the actions. For instance, given the trigger `t(int x) = {Goo g.goo(x)entry}`, one may define the transition `q0 -> q1 [t\x == 8]`, where `x` is the argument of the trigger `t`.

Now, let us illustrate some of the previous expressions with the following example.

```
TRANSITIONS {
  q0 -> q1 [tn \ c == 8 \ v = 0 ; inc(v) ;]
  q0 -> q2 [f \ c == 2 \ ]
  q1 -> q3 [g \ \ v=2 ; \gen(e) ; ]
  q2 -> q2 [e?]
}
```

login-logout = λ `u` : `User`.

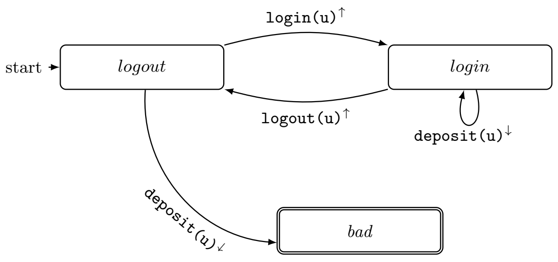


Fig. 3. *ppDATE* template describing properties about the log in and log out of users.

Note that the action of the third transition includes the execution of the action `\gen(e)`. In the input language, this is the action which generates the action event `e`. However, in the theory [5], like it was introduced in Sec. 2, this action is represented with the symbol !, i.e., `e`! generates the action event `e`. The main reason why we decided not to use the same notation in our input language as the one used in the theory is that both JML and Java use the symbol ! as the boolean negation. Thus, we consider that by introducing action `\gen` instead, we are avoiding the possible confusion which may arise in relation to whether `v`! refers to the negation of the boolean variable `v`, or the generation of the action event `v`. In addition, the trigger `e?` in the fourth transition represents the trigger which is activated whenever the action event `e` occurs.

Regarding the use of (Java) programs in the actions, the tool only supports the use of method calls, e.g., `inc(v)` in the first transition. However, it is possible to use the section METHODS (see Sec. 4.6) to define programs as (Java) methods. Then, one simply has to make a method call to them.

4.3 TEMPLATES

In addition to *ppDATE*s which exist up-front, and 'run' from the beginning of a program's execution, new *ppDATE*s can be created by existing ones. For instance, one may want to create a separate 'observer' for each new user logging into a system. For that, one needs to be able to define parameterised *ppDATE*s, which we call *templates*, and allow *ppDATE*s to create new instantiations of them. Fig. 3 illustrates an example of a *ppDATE* template called *login-logout* which, given a user `u`, describes the property "the user has to log in to perform a deposit".

Section `TEMPLATES` lists tagged *ppDATE* templates. Below, we show the syntax of this section.

```
TEMPLATES {
 TEMPLATE id_template (args) {
    VARIABLES { -- definition of the variables -- }
    TRIGGERS  { -- definition of the triggers -- }
    PROPERTY name { -- definition of the property -- }
 }
}
```

Each template is described within a subsection `TEMPLATE`, whose header is followed by a (unique) name `id_template` assigned to the template, and a list of parameters `args` used to generalise the definition of the templates. Note that as a template describes a *ppDATE*, the subsections `VARIABLES`, `TRIGGERS`, and `PROPERTY` are defined just like it is decribed in Sec. 4.2. Below, we illustrate how the *ppDATE* template in Fig. 3 could be written using this syntax.

```
TEMPLATES {
 TEMPLATE login-logout (User u) {
  TRIGGERS  {
    login_exit(String username, int pwd)
        = {User u1.login(username, pwd)exit()} where {u = u1}
    logout_exit() = {User u1.logout()exit()} where {u = u1}
    deposit_entry(int money)
        = {User u1.deposit(money)entry} where {u = u1}
  }
  PROPERTY deposit {
    STATES {
       STARTING { logout ; }
       ACCEPTING { login ; }
       BAD { bad ; }
    }
    TRANSITIONS {
       logout -> login [login_exit]
       logout -> bad [deposit_entry]
       login  -> logout [logout_exit]
       login  -> login [deposit_entry]
    }
  }
 }
}
```

Note that as on the definition of the triggers one has to describe the target object (with its type) on which the methods are called, we have introduced a `where` clause to express that the target object used in such a definition is actually the parameter of the template. This is not the only use for this clause. See Sec. 4.7 for more details about this.

Regarding the instantiation of a template, it is accomplished by using the action `create` on the transition of a *ppDATE*. This action receives as

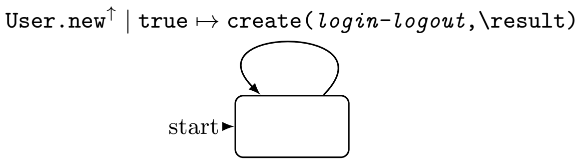


Fig. 4. *ppDATE* in charge of creating instances of the template *login-logout*.

arguments the name of the *ppDATE* template to be instantiated and a list of values to instantiate the parameterised arguments of the template, and it generates the instance of the template. For example, Fig.4 illustrates a *ppDATE* which creates an instance of the template *login-logout* (Fig. 3) upon declaration of an object of class `User`. Here, `\result` represents the (concrete) object of class `User` which was created. In addition, the trigger `User.new`$^{\uparrow}$ is activated when such a creation occurs.

4.4 CINVARIANTS

Section `CINVARIANTS` lists the definitions of class invariants which may need to be considered during the verification of the properties. Its syntax is describided as follows:

```
CINVARIANTS {
 class { invariant }
}
```

Here, `class` represents a Java class in the program under scrutiny whose implementation has to preserve the invariant definition described by `invariant`. Such invariants follow JML-like syntax and pragmatics. Below we illustrate an example of this section.

```
CINVARIANTS {
 Foo { v <= 10 }
 Foo { count >= 0 }
}
```

Note that if no class invariants are needed on a specification, then this section may be omitted. In addition, the actual version of the tool only uses the class invariants during the static verification of the Hoare triples. However, we are currently working to include the verification of class invariants at runtime as well.

4.5 HTRIPLES

Section *HTRIPLES* lists tagged Hoare triples. Its syntax is described as follows:

```
HTRIPLES {
  HT hoare_triple_name {
    PRE { -- precondition -- }
    METHOD { -- method to verify -- }
    POST { -- postcondition -- }
    ASSIGNABLE { -- variables modified -- }
  }
}
```

Each Hoare triple is described within a subsection `HT`, whose header is followed by the name assigned to the Hoare triple. This name is unique for each Hoare triple, and it is used to associate the Hoare triples with the states of a *ppDATE*. Subsection `HT` is composed by four parts: `PRE`, which describes the pre-condition of the Hoare triple; *POST*, which describes the post-condition of the Hoare triple; `METHOD`, which describes which is the method that has to fulfil the Hoare triple; and `ASSIGNABLE`, which lists the variables that might be modified when the method under scrutiny on the Hoare triple is executed. Here, `PRE`, *POST*, and `ASSIGNABLE` follow JML-like syntax and pragmatics. Regarding `METHOD`, its content is an expression of the form `file.method`, where `method` is the name of the method related to the Hoare triple, and `file` is the Java file where the previous method is implemented. Let us illustrate this with following example.

```
HTRIPLES {
  HT add_ok {
    PRE { v }
    METHOD { Foo.inc }
    POST { count == \old(count)+1 }
    ASSIGNABLE { count }
  }
  HT add_err {
    PRE { !v }
    METHOD { Foo.inc }
    POST { count == \old(count) }
    ASSIGNABLE { \nothing }
  }
}
```

Note that this section may contain several subsections `HT`, one per each Hoare triple. In addition, if no Hoare triples are included as part of a *ppDATE* specification, then this section may be omitted.

4.6 METHODS

Section *METHODS* is an optional section which allows to include method declarations as part of a specification. These methods will be included

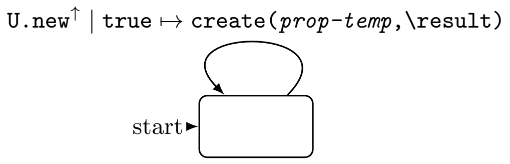


Fig. 5. *ppDATE* in charge of creating instances of the template *prop-temp*.

as part of the implementation of the monitor generated by the tool. Its syntax is described as follows:

```
METHODS {
  type method(arguments) { -- method implementation -- }
}
```

Methods are declared following standard Java notation. However, access modifiers (i.e., *public*, *protected*, *private*) are not necessary when declaring these methods. If a method is declared as *static* method, then monitor variables will not be accessible within that particular method. Below, we illustrate an example of this section.

```
METHODS {
  boolean compare (int x, int y) { return (x == y); }
  int four() { return 4 ; }
}
```

4.7 Extra Features

This section describes some extra features added to the tool and its input language which are not covered by the theory in [5].

PINIT Definition in Section PROPERTY When describing a *ppDATE* specification, it is quite common to have some *ppDATE*s only focus on creating instances of a template upon declaration of an object. Such *ppDATE*s would look like the *ppDATE* illustrated in Fig. 5, which creates an instance of the template prop-temp every time an object of the class U is created. Here, \result represents the (concrete) object of class U which was created. In addition, the trigger U.new$^{\uparrow}$ is activated when such a creation occurs.

Therefore, we decided to include a special subsection PINIT as part of the section PROPERTY, which can be used to specify these kind of *ppDATE*s in a simple manner. The syntax of this subsection is as follows:

```
PROPERTY property_name {
   PINIT { (template_name,Class) }
}
```

where `property_name` is the name of the property (as described in Sec.4.2), `template_name` is the name of a *ppDATE* template defined in section TEMPLATE, and `Class` is the name of the class associated to the declared object. Below, we illustrate how the *ppDATE* from Fig.5 is described using this special subsection.

```
PROPERTY example {
   PINIT { (prop-temp,U) }
}
```

Like in Fig. 5, property `example` describes a *ppDATE* which has a single state with only a loop transition which is fired every time an object of the class `U` is created, leading to an instantiation of the template `prop-temp` using the created object as argument.

Where clause When declaring a trigger, any of its arguments can be bound to a variable which is not directly related to the method arguments. For instance, let us assume that we have to perform some processing on a particular value, and that we want that depending on its result, a *ppDATE* fires a transition (or not). Then, by using a where clause right after a trigger definition one can use one argument of the trigger to as a bound for that particular value. Consider the next example:

```
TRIGGERS {
 goo(boolean y) = {Goo g.foo()entry} where {y = g.IsValid();}
}
```

Here, we do not have any interest in the whole object `g`, but we simply need to know if its a valid object or not, fact which can be computed using the method `IsValid()`, in order to send the *ppDATE* to either the state `q2`, or `bad`, respectively. Then, one can use the boolean argument `y` of the trigger for binding the result of that method. This would allow us to write transitions like the following ones:

```
TRANSITIONS {
  q1 -> q2  [goo\ y]
  q1 -> bad [goo\ !y]
}
```

Here, remember that it is not necessary to write the arguments of a trigger in the trigger component of a transition, but one can refer to them in both the conditions and the actions.

Furthermore, any variable which is not directly bound to the method arguments is initialized in the where clause. This is done by checking that there is at least one assignment statement with the unbound variable on the left-hand side.

Note that the statements in the where clause can be any valid JAVA statements and these can call any relevant method from imported packages, and that the use of curly brackets in this clause is compulsory.

Foreach Construct The `FOREACH` construct can be used as a simplistic alternative to the use of *ppDATE* templates. Consider the following *ppDATE*:

```
GLOBAL {
  TRIGGERS {
    log(User user) = {Interface f.login(User user)entry}
    out(User user) = {Interface f.logout(User user)entry}
  }
  PROPERTY example {
    STATES {
      ACCEPTING { logout ; }
      BAD { bad ; }
      STARTING { login ; }
    }
    TRANSITIONS {
       logout -> login [log\\create(deposit-temp,user)]
       logout -> bad [out]
       login  -> logout [out]
       login  -> bad [log]
    }
  }
}

TEMPLATES {
 TEMPLATE deposit-temp (User u) {
  TRIGGERS {
   dep(int val) = {User u1.deposit(val)entry} where {u = u1;}
  }
  PROPERTY deposit { --- }
 }
}
```

On this *ppDATE*, every time a user logs in the interface, an instance of the template `deposit-temp` is created in order to runtime verify the property `deposit` for that user.

Now, let us introduce a similar *ppDATE* to the one described above, but written using the foreach construct:

```
GLOBAL {
  TRIGGERS {
    log(User user) = {Interface f.login(User user)entry}
    out(User user) = {Interface f.logout(User user)entry}
  }
  PROPERTY example {
    STATES {
      ACCEPTING { logout ; }
      BAD { bad ; }
      STARTING { login ; }
    }
    TRANSITIONS {
       logout -> login [log]
       logout -> bad [out]
       login  -> logout [out]
       login  -> bad [log]
    }
  }

  FOREACH (User u) {
   TRIGGERS {
    dep(int val) = {User u1.deposit(val)entry} where {u = u1;}
   }
   PROPERTY deposit { --- }
  }
}
```

In this version of the *ppDATE*, as soon as an object of the class `User` is created, a *ppDATE* verifying the property `deposit` is generated. Here, remember that this is not what happen the template version of the *ppDATE*, where the *ppDATE* verifying property `deposit` is only created when a user logs in.

Using a foreach construct may seem simpler than using a *ppDATE* template. However, one have to consider the following points when using it: (i) this construct introduces a context to the *ppDATE*, i.e., the triggers, variables and transitions will now be in a particular context. Hence, each trigger should specify its context so that the *ppDATE* which will be affected will only be the one belonging to that particular context. This is done by using the where clause associated to each trigger; (ii) *ppDATE*s for verifying the properties within a foreach are always going to be generated upon creation of an object, even if the execution of the program does not require to verify them; (iii) *ppDATE* templates are much more expressive than this construct.

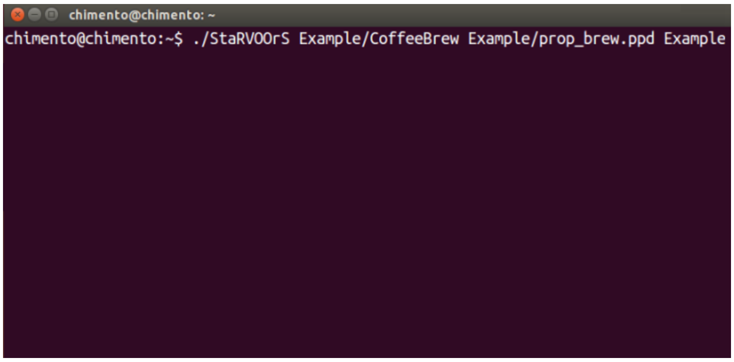
5 Using StaRVOOrS

In this section we depict how StaRVOOrS works by running the tool on the coffee machine example described in Sec. 2. Both the *ppDATE* specification written in the input language of the tool, and a simplistic implementation of the coffee machine system, together with two big case studies based on Mondex [1] and SoftSlate (a real Java cart application) [2], can be found in [3], under the section *Downloads*. In addition, we have included the content of the files of the running example in the appendix A.

5.1 Running StaRVOOrS

In order to run StaRVOOrS, as it is illustrated in Fig. 6, the following input should be provided:

(i) the path to the main directory of the Java files to be verified, e.g., `Example/CoffeeBrew`.
(ii) a *ppDATE* specification for the provided program written in the input language of the tool, e.g., `Example/prop_brew.ppd`.
(iii) the path of the output directory where the files generated by the tool are going to be placed, e.g., `Example`.



```
chimento@chimento: ~
chimento@chimento:~$ ./StaRVOOrS Example/CoffeeBrew Example/prop_brew.ppd Example
```

Fig. 6. Runnig StaRVOOrS

5.2 StaRVOOrS ouput

Fig.7 illustrates all the files generated by StaRVOOrS when it is used to analise the running example. This output consists of: the monitor files generated by Larva (folder `aspects` and folder `larva`), the files generated by StaRVOOrS to runtime verify partially proven Hoare triples (folder `ppArtifacts`), an instrumented version of the source code (folder `CoffeMachine`), the xml file used by StaRVOOrS to optimise the *ppDATE* specification (*out.xml*), a report explaining the content of the .xml file (*report.txt*), the optimised version (if any) of the provided *ppDATE* specification (*prop_brew_optimised.ppd*), and the *DATE* specification obtained as a result of translating the (optimised) *ppDATE* (*prop_brew.lrv*). Note that StaRVOOrS does not modify the provided source code, it creates an instrumented version of it. Thus, at the time of monitoring the code, the instrumented version of the source code is the one which should be used.

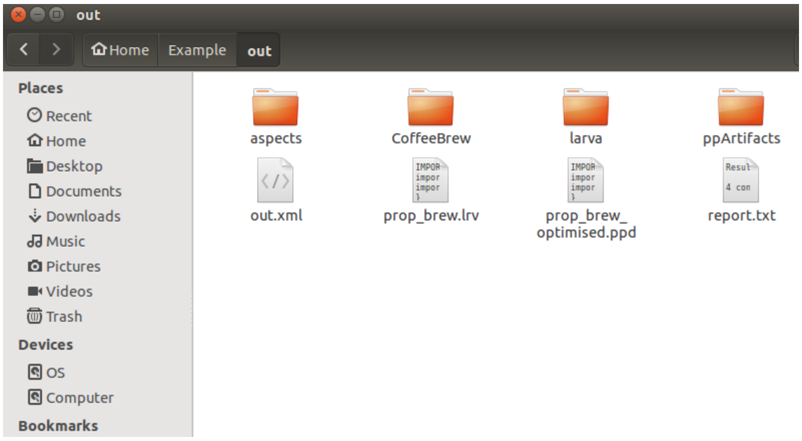


Fig. 7. StaRVOOrS output

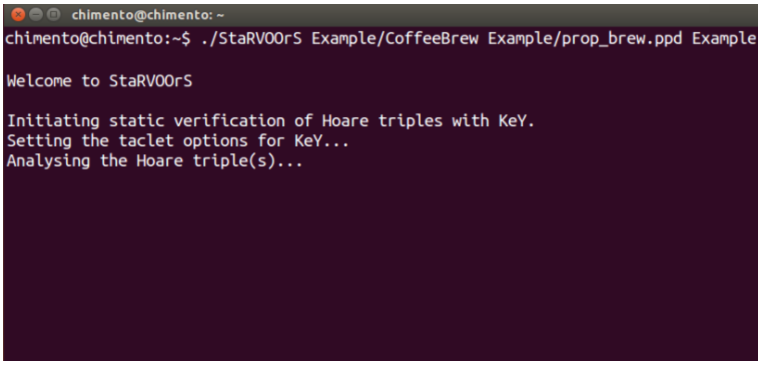
5.3 StaRVOOrS execution insights

StaRVOOrS is a fully automated tool. However, in order to have a better understanding of its execution, below we will explain it in three stages. Note that during each one of this stages, StaRVOOrS will produced some output on the terminal. We will illustrate such an output through figures.

The first stage corresponds to the static verification of the Hoare triples using KeY. Fig. 8 shows the output produced by the tool on the terminal during this stage. At first, KeY (taclet) options are set. This options are parameters which, for instance, indicate KeY which rules of its sequent calculus it is able to use during the verification of a property. For the time being, we are just using the standard options. Then, KeY is ran.

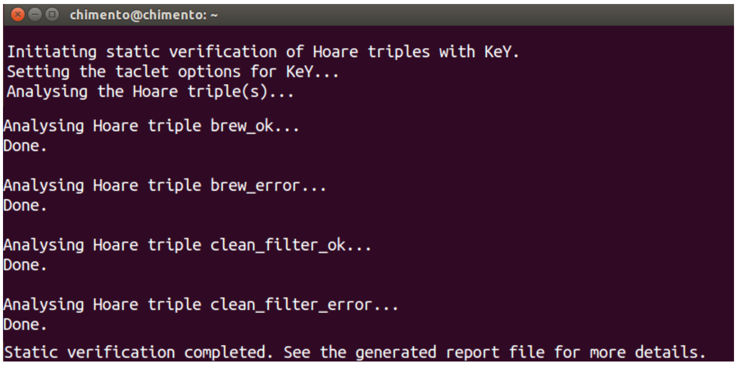
While KeY analyses all the Hoare triples, every time a proof attempt is saturated, some information related to this analysis is given as output in the terminal. Fig. 9 illustrates this. Once that KeY is done verifying all the Hoare triples, it generates a *out.xml* describing its results. This file is not intended for the user, it is used by StaRVOOrS to optimise the *ppDATE* specification for runtime checking. However, in order to give to the user some understandable feedback about what happened during the static verification of the contracts, StaRVOOrS generates a file *report.txt* which briefly explains the content of the .xml file.

The second stage corresponds to the refinement of the specification. In this stage, all the Hoare triples which were fully proven are removed from the *ppDATE*, and those which were only partially proven are modified by strengthening their pre-conditions including the conditions which lead



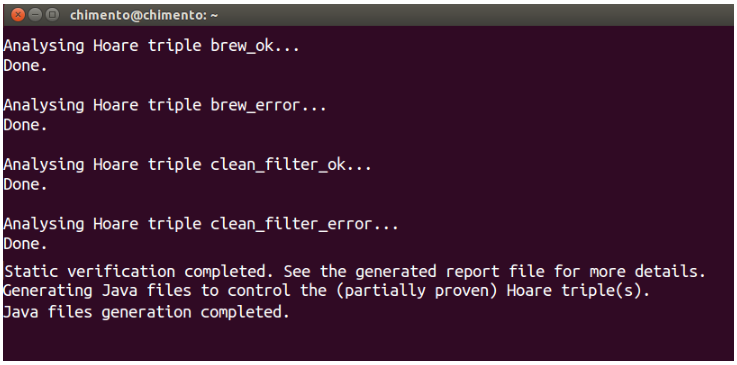
```
chimento@chimento: ~
chimento@chimento:~$ ./StaRVOOrS Example/CoffeeBrew Example/prop_brew.ppd Example

Welcome to StaRVOOrS

Initiating static verification of Hoare triples with KeY.
Setting the taclet options for KeY...
Analysing the Hoare triple(s)...
```

Fig. 8. Initiating Static Verification



```
chimento@chimento: ~
 Initiating static verification of Hoare triples with KeY.
 Setting the taclet options for KeY...
 Analysing the Hoare triple(s)...

Analysing Hoare triple brew_ok...
Done.

Analysing Hoare triple brew_error...
Done.

Analysing Hoare triple clean_filter_ok...
Done.

Analysing Hoare triple clean_filter_error...
Done.
Static verification completed. See the generated report file for more details.
```

Fig. 9. Output shown on the terminal during static verification



```
chimento@chimento: ~
Analysing Hoare triple brew_ok...
Done.

Analysing Hoare triple brew_error...
Done.

Analysing Hoare triple clean_filter_ok...
Done.

Analysing Hoare triple clean_filter_error...
Done.
Static verification completed. See the generated report file for more details.
Generating Java files to control the (partially proven) Hoare triple(s).
Java files generation completed.
```

Fig. 10. Optimization and files generation after static verification

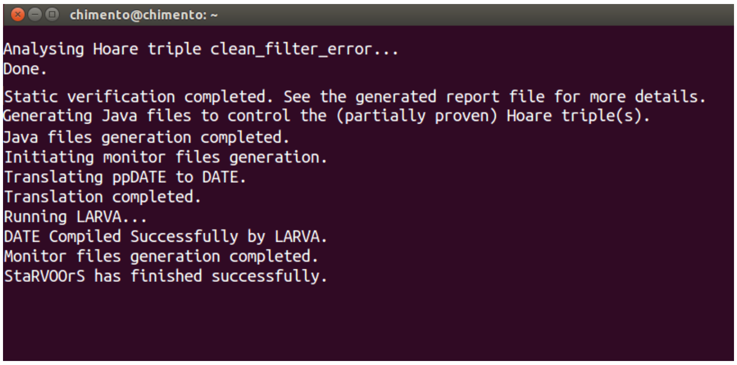
to an unclosed path on a proof. For instance, as it can be read on the report file (see appendix B), the pre-condition of the Hoare triple *brew_ok* is strengthen with the addition of the condition `active == TRUE`.

Whenever it is necessary to runtime verify partially proven Hoare triples, StaRVOOrS instruments the provided Java files by adding a new parameter to the method(s) associated to the Hoare triple(s). This new parameter is used to distinguish different calls to the same method. This change is introduced in the refined *ppDATE* specification as well (see appendix B). Besides, StaRVOOrS generates several files within folder the `ppArtifacts` which are used to runtime verify the Hoare triples. For instance, the file *HoareTriplesPPD.java* contains the implementation of the methods which are used to verify the pre- and post-conditions of the Hoare triples. In addition, file *IdPPD.java* will be used to generate the value of the new parameter added to the methods. Once this stage is over, the terminal will look like Fig. 10.

The third stage corresponds to the generation of the runtime monitor. In order to do so, the refined *ppDATE* specification is translated by StaRVOOrS to a *DATE* specification (file *prop_brew.lrv* in our running example). Then, Larva is used to generate the monitor files from the *DATE*. After the execution of Larva is completed, leading to the generation of the files in the folders `aspects` and `larva`, StaRVOOrS execution is completed as well. The terminal will reflect this, as it is illustrated in Fig. 11.

5.4 Running the application with the generated monitor

To run the (generated) instrumented version of the program, let us call it P, together with the monitor, one can generated an executable jar file, and then run it on a Java virtual machine. We will use Java 1.7 to compile



```
chimento@chimento: ~
Analysing Hoare triple clean_filter_error...
Done.
Static verification completed. See the generated report file for more details.
Generating Java files to control the (partially proven) Hoare triple(s).
Java files generation completed.
Initiating monitor files generation.
Translating ppDATE to DATE.
Translation completed.
Running LARVA...
DATE Compiled Successfully by LARVA.
Monitor files generation completed.
StaRVOOrS has finished successfully.
```

Fig. 11. Monitor Generation

the Java files. However, due to compatibility issues with Larva, when compiling the aspects one has to use the version 1.5. Note that the aspects have to be compiled using an AspectJ compiler. We recommend the *ajc* compiler. In addition, to run the jar file one has to use *aj5* (like command *java*, but with support for AspectJ), or similar. Below, we provide a short script explaining how to create such a jar file.

First, go inside the output directory.

cd Example/out

Second, copy the folders `larva` and `ppArtifacts` into the main folder of *P*.

cp -r larva CoffeeBrew
cp -r ppArtifacts CoffeeBrew

Third, create a directory named `Build`, and compile *P* using the option *-target 1.7* in such a way that the compiled files are placed within `Build`.

mkdir Build
*javac -target 1.7 $(find CoffeeBrew -name *.java) -d Build*

Next, create an executable jar file from the files in `Build`.

jar cfe coffeeM.jar main.CMachine -C Build .

Now, one has to weave the aspects into the jar file. In order to do so, the files in folder aspects have to be compiled using an AspectJ compiler. Here, we use *ajc*. Note that it is usually recommend to generate a new jar file when the aspects are weaved.

ajc -1.5 -sourceroots aspects/ -inpath coffeeM.jar -outjar coffeeM_asp.jar

Finally, this weaved executable jar file corresponds to the compilation of *P* together with the monitor. Therefore, running it would mean the one is running a monitored version of *P*. To execute the weaved jar file one can use *aj5* as follows:

aj5 -jar coffeeM_asp.jar

References

1. MasterCard International Inc. Mondex. `www.mondexusa.com/`.
2. SoftSlate Commerce. `www.softslate.com/`.
3. StaRVOOrS web page. `cse-212294.cse.chalmers.se/starvoors`.
4. Wolfgang Ahrendt, Bernhard Beckert, Richard Bubel, Reiner Hähnle, Peter H. Schmitt, and Mattias Ulbrich, editors. *Deductive Software Verification—The KeY Book*, volume 10001 of *LNCS*. Springer, 2016. to appear.
5. Wolfgang Ahrendt, Jesús Mauricio Chimento, Gordon J. Pace, and Gerardo Schneider. A Specification Language for Static and Runtime Verification of Data and Control Properties. In *FM'15*, volume 9109 of *LNCS*. Springer, 2015.
6. Christian Colombo, Gordon J. Pace, and Gerardo Schneider. Dynamic Event-Based Runtime Monitoring of Real-Time and Contextual Properties. In *FMICS'08*, volume 5596 of *LNCS*, pages 135–149. Springer-Verlag, September 2009.
7. Christian Colombo, Gordon J. Pace, and Gerardo Schneider. LARVA - A Tool for Runtime Monitoring of Java Programs. In *SEFM'09*, pages 33–37. IEEE Computer Society, 2009.
8. David Harel, Dexter C. Kozen, and Jerzy Tiuryn. *Dynamic logic*. Foundations of computing. the MIT Press, Cambridge (Mass.), London, 2000.
9. Gary T. Leavens, Erik Poll, Curtis Clifton, Yoonsik Cheon, Clyde Ruby, David Cok, Peter Müller, Joseph Kiniry, and Patrice Chalin. *JML Reference Manual. Draft 1.200*, 2007.

A Running Example Files

A.1 Coffee Machine Implementation

```
public class CMachine {
   public int cups;
   public int limit;
   public boolean active;

   CMachine(int limit) {
        this.limit = limit;
        cups = 0;
        active = false;
   }

   public void cleanF() {
       if (!active)
          cups = 0;
   }

   public void brew() {
       if (!active && cups < limit)
          cups++;
   }
}
```

A.2 *ppDATE* Specification for the Coffee Machine

```
IMPORTS { import main.CMachine; }

GLOBAL {
 EVENTS {
   brew_entry() = {CMachine cm.brew()}
   brew_exit() = {CMachine cm.brew()uponReturning()}
   cleanF_entry() = {CMachine cm.cleanF()}
 }
 PROPERTY prop {
  STATES {
    BAD { bad ; }
    NORMAL { q2 (brew_error,clean_filter_error) ;}
    STARTING { q (brew_ok,clean_filter_ok) ; }
  }
  TRANSITIONS {
    q -> q2 [brew_entry \ cm.cups < cm.limit ]
    q2 -> q [brew_exit ]
    q2 -> bad [ brew_entry ]
    q2 -> bad [ cleanF_entry ] }
 }
}

HTRIPLES {
 HT brew_ok {
   PRE {cups < limit}
   METHOD {CMachine.brew}
   POST {cups == \old(cups)+1}
   ASSIGNABLE {cups} }
 HT brew_error {
   PRE {cups < limit}
   METHOD {CMachine.brew}
   POST {cups == \old(cups)}
   ASSIGNABLE {cups} }
 HT clean_filter_ok {
   PRE {true}
   METHOD {CMachine.cleanF}
   POST {cups == 0}
   ASSIGNABLE {cups} }
 HT clean_filter_error {
   PRE {true}
   METHOD {CMachine.cleanF}
   POST {cups == \old(cups)}
   ASSIGNABLE {cups} }
}
```

B Files Produced by StaRVOOrS

B.1 report.txt

Results of the Static Verification of Hoare triples

4 contract(s) were analysed:

* No contract(s) were fully proved.

* 4 contract(s) were partially proved:
clean_filter_error ⟶ New condition added to its precondition is
!(active == true)
clean_filter_ok ⟶ New condition added to its precondition is
active == true
brew_error ⟶ New condition added to its precondition is
!(active == true)
brew_ok ⟶ New condition added to its precondition is
active == true

B.2 Refined Version of the *ppDATE* Specification for the Coffee Machine

```
IMPORTS { import main.CMachine; }

GLOBAL {
 TRIGGERS {
  brew_entry(Integer id) = {CMachine cm.brewAux(id)entry}
  brew_exit(Integer id) = {CMachine cm.brewAux(id)exit()}
  cleanF_entry(Integer id) = {CMachine cm.cleanFAux(id)entry}
  cleanF_ex(Integer id) = {CMachine cv.cleanFAux(id)exit()}
 }
 PROPERTY prop {
  STATES {
    BAD { bad ; }
    NORMAL { q2 (brew_error,clean_filter_error) ;}
    STARTING { q (brew_ok,clean_filter_ok) ; }
  }
  TRANSITIONS {
    q -> q2 [brew_entry \ cm.cups < cm.limit ]
    q2 -> q [brew_exit ]
    q2 -> bad [ brew_entry ]
    q2 -> bad [ cleanF_entry ] }
 }
}

HTRIPLES {
 HT brew_ok {
   PRE {cups < limit && active== true}
   METHOD {CMachine.brew}
   POST {cups == \old(cups)+1}
   ASSIGNABLE {cups} }
 HT brew_error {
   PRE {cups < limit && !(active== true)}
   METHOD {CMachine.brew}
   POST {cups == \old(cups)}
   ASSIGNABLE {cups} }
 HT clean_filter_ok {
   PRE {true && active== true}
   METHOD {CMachine.cleanF}
   POST {cups == 0}
   ASSIGNABLE {cups} }
 HT clean_filter_error {
   PRE {true && !(active== true)}
   METHOD {CMachine.cleanF}
   POST {cups == \old(cups)}
   ASSIGNABLE {cups} }
}
```