CLARVA: Model-based Residual Verification of Java Programs

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Keywords: Verification, Model-based Analysis, Residual Analysis, Static Analysis.

Abstract: Runtime verification (RV) is an established approach that utilises monitors synthesized from a property language (e.g. temporal logics or some form of automata) to observe program behaviour at runtime, determining compliance of the program with the property at runtime. An issue with RV is that it introduces overheads at runtime, while identifying a violation at runtime may be too late. This can be tackled by introducing light analyses that attempt to prove parts of the property with respect to the program, leaving a *residual* property that induces a smaller monitoring footprint at runtime and encodes some static guarantees. In this paper we present CLARVA as a tool developed for this end for the RV tool LARVA. CLARVA transforms Java code into an automaton-based model, and allows for the incorporation of control-flow analyses that analyse this model against *Dynamic Automata with Timers and Events* or DATES (the property language used by LARVA) to produce residuals that produce an equivalent judgement at runtime.

1 INTRODUCTION

Verification methods can be applied on an program before or after its deployment. Ideally verification is done pre-deployment to ensure a misbehaving program is not deployed, however precise formal verification before deployment can be expensive while parts of the program may be dynamic or unknown pre-deployment. Runtime verification (RV) (Leucker and Schallhart, 2009) is one approach to performing post-deployment verification by instrumenting an application with a monitor that attempts to give a verdict about the program's compliance by analysing an execution prefix. However using only RV means no static guarantees about the program can be given, while RV introduces time and memory overheads at runtime when performed synchronously with the program. In this paper we present a tool addressing both of these problems for LARVA, an established tool for the runtime verification of Java programs.

The LARVA (Colombo et al., 2009) approach expects specifications in the form of *dynamic automata with timers and events*, or DATEs. This language allows the developer to identify a set of program events using AspectJ pointcuts (Kiczales et al., 2001), and an automaton that captures the prohibited event traces. These events are symbolic with parameters that are only bound at runtime (Havelund et al., 2018). Transitions are triggered by such events, while they are guarded by a condition on the program and monitor variable state while possibly acting on the monitor state (e.g. the monitor may keep a counter of the times a certain method is called, prohibiting a certain number of subsequent calls). Moreover DATEs allow for time-triggered events, and also for communication between multiple DATEs through channels, making them more succinct and modular than simple finite-state automata.

The runtime overheads induced by the LARVA approach is thus proportional to the number of times DATE transitions are triggered at runtime and to the expense of executing transition guards and actions. There are several options for reducing runtime overheads then:

(i) reducing the number of times a program triggers an event;

(ii) reducing the number of DATE transitions; and

(iii) reducing or optimising DATE transition guards and actions.

Previous work considers sound approaches to performing these reductions (Azzopardi et al., 2017), with *residual analyses* ensuring that the reduced monitor observations are equivalent to those of the original, at least with respect to the program in question.

In this paper we introduce CLARVA (Azzopardi, 2019), a tool that is able to incorporate analyses that

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effect such sound reductions based on a control-flow model of a Java program. CLARVA is thus intended to be used for relatively inexpensive pre-deployment analyses that produce a *residual* of a DATE and a program with *residual instrumentation* that can be prepared for deployment by LARVA. This was inspired by CLARA (Bodden et al., 2010), a tool for similar analyses but limited to finite-state automata and reductions of instrumentation.

In Section 2 we discuss briefly the background and theoretical framework underpinning CLARVA, while in Section 3 we discuss the architecture and process of the tool. We illustrate the tool using a case study in Section 4, discuss related work in Section 5, and conclude in Section 6.

2 BACKGROUND

In this section we discuss some of the theoretical and practical background behind CLARVA, introducing briefly LARVA (Colombo et al., 2009), how we model programs, and a brief description of residual analyses.

2.1 The LARVA Approach

Common languages for specifications in the RV community can largely be classified as either logic- or automata- based, with automata being also commonly used behind the implementation of logic-based approaches, or as event- or state-based (Falcone et al., 2018). LARVA falls under the event-based approaches that use automata directly as the main specification language. Figure. 1 illustrates a high-level view of the process LARVA undertakes to produced a monitored program. Here we discuss how LARVA handles monitoring for program events and how the specification logic is represented using *dynamic automata with events and timers* (DATEs). Listing. 1 is an example of the LARVA specification language, which we use throughout this section.

2.1.1 Program Events

LARVA uses *AspectJ pointcuts* to allow the specification engineer flexibility in specifying program events of interest independent of the program implementation. These event definitions currently correspond to Java method calls, with different matching modalities (e.g. before the call, after, upon returning, or upon throwing). For example, the LARVA event declaration on lines 7-9 in Listing. 1 matches any return from a call to a method named `login` and that has an integer `id` parameter.

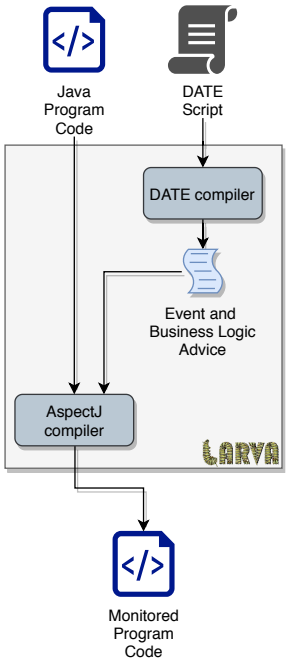


Figure 1: LARVA process.

```
1 GLOBAL{
2  VARIABLES{
3   int limit = 250;
4  }
5  FOREACH(User u){
6   EVENTS{
7    login(User u1)
8              = {*.login(int id)uponReturning
                  ()}
9      where {u1 = users.get(id); u = u1;}
10    ...
11   }
12   PROPERTY userBehaviour {
13    STATES{
14     ACCEPTING {good}
15     BAD {bad}
16     NORMAL {loggedIn}
17     STARTING {loggedOut}
18    }
19    TRANSITIONS{
20     loggedIn->loggedOut[login(u1)\u1.
           activated\]
21     ...
22    }}}}
```

Listing 1: Partial view of a LARVA script.

The left-hand side of this declaration represents the event as will be used in the DATE. Note how this event is symbolic in that it is instantiated at runtime with some program data (see Line 9). The right-hand side syntax is close to that used by AspectJ for pointcuts. The syntax also allows for a single DATE event (left-hand declaration) to correspond to multiple program events.

2.1.2 DATEs

Finite-state automata are traditionally parametrised by a finite set of events, with transitions tagged by these events and the automaton representing the acceptable set of event traces. As discussed in the previous section, events here are instead *symbolic* in that they may have attached certain information about the program. The event set here effectively is not finite — note how for arbitrary $n \in \mathbb{N}$ then `login(n)` is an individual event. The automata used by LARVA, DATEs, are then forms of *symbolic automata* that are able to deal with this infinite set of events with a finite set of states and transitions. Moreover DATEs allow us to maintain some internal monitoring variable state, effectively allowing for a more succinct representation than finite-state automaton.

A DATE transition is of the form: $q- > q'\ [e \backslash c \backslash a]$, where q and q' are automata states, e is a symbolic event, c is a predicate on the event parameters (which are bound at runtime) and the monitoring variable state, and a is a transformation of the monitoring variable state. For example, Line 20 in Listing. 1 defines a transition that allows a user to login only when the user account's `activated` attribute is true.

2.1.3 Typestate

A concern in RV is that of *typestate*, where a monitor is intended to be replicated for each object of a certain type. For example, given a `Stream` object one may want to ensure that each such object is not used after it is closed. In this case typestate logic is used to ensure that a monitor is instantiated for each such object at runtime, where the monitor only listens to events of the object they are associated with. This functionality is also present for DATEs, through the `FOREACH` construct. In Listing. 1 Line 5 illustrates how the property defined by the script is parametrised by a `User` typestate, with each event identifying the relevant user (see Line 9).

2.2 Modeling Programs

DATEs abstract away from program details through DATE events. In the same manner when performing static analysis of a program it is ideal to work at this level of abstraction rather than at the level of program code. Moreover, since our specifications are automata-based, representing programs in an automata-based formalism allows us to exploit the rich theory behind automata in performing verification and analysis. In CLARVA we choose to internally represent Java methods as non-deterministic automata with states representing Java statements and transitions tagged with DATE events. A program is thus represented as a set of such automata.

These automata in effect represent an over-approximation of the possible execution traces of the program. This is only sound and not complete because we are ignoring program data-flow. In Section 4 we illustrate an example of such automata and their corresponding Java methods.

2.3 Residual Analysis

CLARVA is essentially a tool for what we call *residual analysis*. Here we give a brief semi-formal introduction to this approach.

The verification problem is essentially whether a program P satisfies a property π, which we denote by $P \vdash \pi$. Residual analysis, given a program P, is concerned with the reduction of the property π to a property π' such that π and π' are interchangeable for the purposes of verification, i.e.

$$P \vdash \pi \iff P \vdash \pi'.$$

This reduction does not occur blindly but by considering the known properties of the system, in effect in the form of a *model* of the system (Azzopardi et al., 2016). Then, a residual operation can take into account this model when producing a residual property. This can be represented as follows, where the residual operation is denoted by the symbol $\backslash$:

$$P \vdash M \implies (P \vdash \pi \backslash M \iff P \vdash \pi).$$

This notion allows us to use $\pi \backslash M$ instead of π for the purposes of verification, with the guarantee that the result is equivalent to verifying π, for any program satisfying M. One question is how to acquire the model M. An option is through analysis (which we do in CLARVA by creating an automaton representation of the program), while in other cases the program P may be synthesized from an existing model M.

This condition is however too wide for the purposes of RV. Consider that RV is concerned with the compliance of single execution traces, while our verification operator here ($\vdash$) is concerned with the compliance of the whole program. Then, the residual condition here does not guarantee that an event trace of P will be given the same verdict by $\pi \backslash M$ as by π. Instead we can refine this by considering the event traces (with events from an alphabet A[1]) induced by a program using an operator: $T(P) : 2^{A^*}$, and a trace

[1] Note that here A represents a set of events. These events can also be symbolic, e.g. the events may be associated with a snapshot of the program variable state.

compliance operator $\Vdash$[2]:

$$P \vdash M \implies (\forall t \in T(P) \cdot t \Vdash \pi \setminus M \iff t \Vdash \pi).$$

This correctness condition is what we require out of any property residual analysis included in CLARVA, to ensure any reduced property can be used for RV without any unexpected behaviour.

The overheads of RV are not associated only with the property being monitored, but also with the amount of instrumentation inserted into the program to trigger property events. To model this consider that a program P can be represented by a set of executions (where STMT is the type of statements) $Ex \subseteq 2^{\mathsf{STMT}^*}$, and then instrumentation can be modeled as a function from executions to event traces $inst_P : \mathsf{STMT}^* \mapsto A^*$. Finally an instrumentation reduction of a program with respect to a property, i.e. the function $inst_{P\setminus\pi}$, is correct only if the event trace induced by an execution prefix by both instrumentation functions is given the same verdict by π (using *pre* for a function that gives all the prefixes of a set of traces):

$$\forall e \in pre(Ex) \cdot inst_P(e) \Vdash \pi \iff inst_{P\setminus\pi}(e) \Vdash \pi.$$

This correctness condition for instrumentation reductions finalises our brief treatment of the residual analysis CLARVA incorporates. For a more in depth treatment see (Azzopardi et al., 2017).

3 ARCHITECTURE

The CLARVA process is made up of three different phases (Figure. 2): (i) the abstraction phase where the Java code and LARVA script are abstracted into automata models; (ii) the residual analysis phase where different residual analyses are applied in succession; and (iii) the concretisation phase where the internal representations are transformed into executable format.

During the abstraction phase the input program and LARVA script are abstracted respectively by a control-flow model (as described in Section 2.2) and a DATE (as described in Section 2.1.2). In the abstraction of programs the Java bytecode analyser Soot (Vallée-Rai et al., 1999) is used to analyse the program code and structure to create appropriate automata models of the program methods while the DATE events are used to instrument these automata with events. The resulting automata represent an over-approximation of the execution traces of the program at runtime with respect to the given DATE.

[2]We require the condition that a program is compliant with a property iff all its traces are also compliant with the property: $P \vdash \pi \iff \forall t \in T(P) \cdot t \Vdash \pi$.

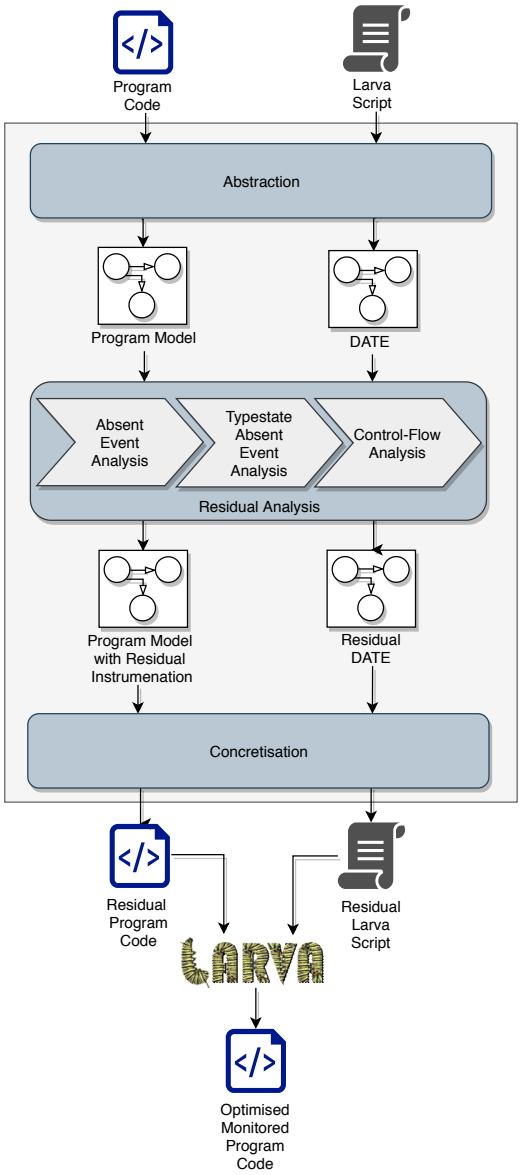


Figure 2: CLARVA process.

The residual analysis phase incorporates a number of residual analyses, as described in (Azzopardi et al., 2017), obeying the conditions defined in Section 2.3. These analyses compare the program model with the DATE to produce reduced version of each. For example, *absent event analysis* analyses the model to identify events used by the DATE but not actually triggered by the program and reducing the DATE appropriately. The subsequent analyses introduce more precision, e.g. *typestate absent event analysis* projects the same analysis for every possible runtime typestate object, and *control-flow analysis* uses composition of program and property automata to identify transitions in the DATE that are never activated for any typestate.

Dually these can be used to identify instrumentation points in the program that can be silenced (i.e. such that they do not trigger events at runtime). These analyses are subsequently more precise, using imprecise and cheaper analyses to possibly reduce the expense of the next stage. These analyses are performed multiple times until a fix-point is reached. New residual analyses can be introduced at this stage to create finer residuals.

The concretisation phase is the dual of the abstraction phase, using the internal abstractions to output a Java program with optimised instrumentation and an optimised monitor. The instrumentation identified as unnecessary by the residual analysis phase is used to transform the original Java code such that function calls triggering events unnecessarily do not trigger events after LARVA instrumentation. This transformation does not change the program behaviour although it does introduce a new Java class through which calls which should be instrumented pass through. The residual LARVA script instruments methods in this class instead of the original methods, while the monitoring logic may also be reduced, when the analysis phase proves the program cannot explore parts of the DATE, and those parts of the DATE can be pruned.

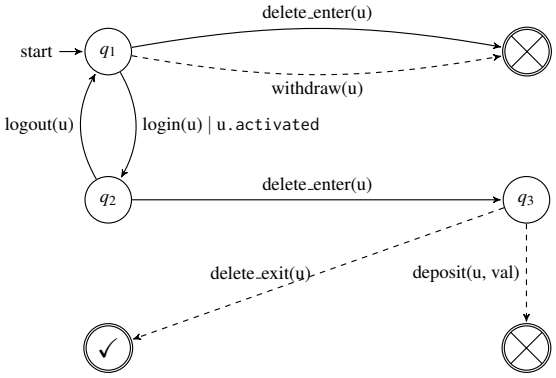
4 CASE STUDY

Here we consider a case study to illustrate the CLARVA approach. Listing. 2 illustrates the code of a simple transaction menu that allows a user to login, logout, transact, deposit, and delete their account, while upon the user requesting deletion, their account is removed from the list of users. Figure. 3 illustrates a specification for this program, where if the user is logged out they are not allowed to delete account or withdraw any funds, while they have a transaction limit when logged in. Moreover during the process of user deletion no deposits into the user's account are allowed.

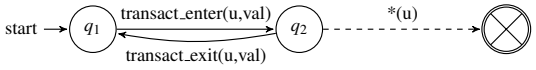
CLARVA processes the program Listing. 2 to produce automata for each method similar to those illustrated in Figure. 4. For illustration purposes these automata are more concise than those actually produced by CLARVA, however they are equivalent language-wise. Note also how the states of these automata are annotated by the relevant program statement.

By analysing the DATE in Figure. 3a against the models in Figure. 4 CLARVA's analyses manage to:

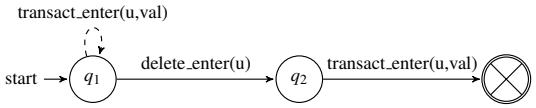
(i) identify that the *withdraw* event is never actually triggerable in the methods (note how it is never called) and thus the *withdraw* transition in the DATE can be safely removed;



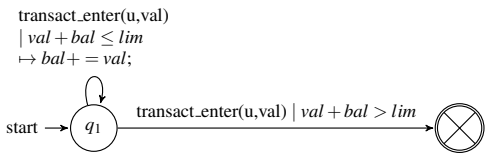
(a) *Deletion and withdrawal only allowed after* `login` *of an activated user, while no deposits are permitted during deletion.*



(b) `transact` *is atomic.*



(c) *No* `transact`*ions allowed after* `delete` *started.*



(d) `transact` *respects limits.*

Figure 3: Several properties expected out of Listing. 2, with dashed transitions representing transitions removed after residual analysis.

(ii) that the *delete* method does not call the *deposit* method and thus the respective transition can be removed from the DATE;

(iii) since the *deposit* event is no longer used in the residual DATE, the *deposit* event instrumented in the *menu* method can be silenced (see Figure. 4a); and

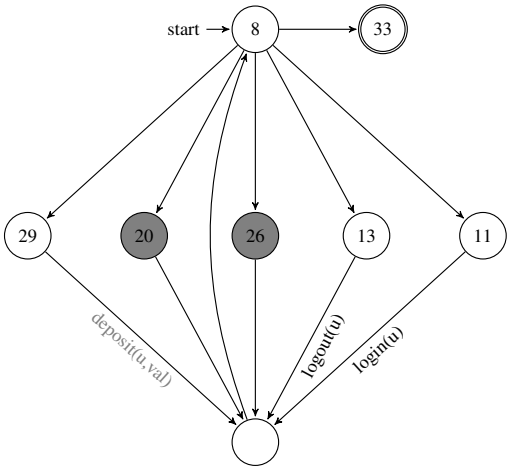
(iv) by analysing the DATE CLARVA identifies that the *delete_exit* event does not need to be monitored, since the residual DATE cannot violate when at state q_3 and thus we can remove the respective DATE transition and event instrumentation in the program. Similarly for Figure. 3b we can determine that no other event occurs during a transaction.

This purely control-flow analysis however fails for properties that require information about the program's variable state after an event occurs, e.g. to

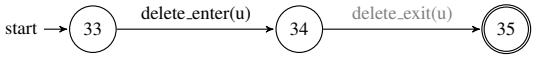
```
1 Map<Integer, User> users;
2
3 public void menu(){
4  int option;
5  User u;
6  boolean open = true;
7  ...
8  while(open){
9   switch(option) {
10      case 0: int id = /* get from input */;
11              u = login(id);
12              break;
13      case 1: if(u != null) logout(u.id);
14               break;
15      case 2: if(u != null){
16                float val = /* get from input
                      */;
17                int to = /* get  from input
                      */;
18                int from = u.id;
19                if(u.balance >= val)
20                    transact(from, to, val);
21              }
22              break;
23      case 3: if(u != null){
24                float val = /* get from input
                      */;
25                int to = u.id;
26                deposit(to, val);
27              }
28              break;
29      case 4: if(u != null) delete(u);
30              break;
31      case 5: open = false;
32    }
33   }
34 }
35
36 public void delete(User u){
37   users.remove(u);
38 }
39
40 public void transact(int from, int to,
41   double val){
42   if(users.get(from).bal >= val
43       && users.containsKey(to)){
44     users.get(from).bal -= val;
45     users.get(to).bal += val;
46   }
47 }
```

Listing 2: Example Java Program.

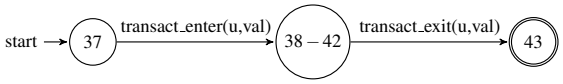
prove Figure. 3c we need to be able to determine that after a `delete` the deleted user is no longer present in the system and thus can no longer be acted upon. However we can optimise the DATE by removing the looping transition on the initial state, since it does not have any effect. Another challenge is the variable state of the DATE, e.g. consider Figure. 3d where transitioning in the DATE depends on a `bal` variable maintained by the DATE.



(a) `menu` *method, with states* 20 *and* 26 *being call states, respectively calling the transact and delete methods.*



(b) `delete` *method.*



(c) `transact` *method.*

Figure 4: Non-deterministic models of Listing. 2 methods, with shaded states representing call states, and faded events representing instrumentation silenced after analysis.

Table 1: Evaluation data of Figure. 3 monitored with and without residual analysis, compared to without monitoring.

Users No.	**Original** *seconds*	**Monitored** *% of orig.*	**Residual**. *% of orig.*	**Savings** *% of mon.*
500	30.86s	313.12%	261.31%	16.55%
600	35.11s	335.06%%	266.56%	20.44
700	40.98s	339.85%	262.57%	22.74%
800	47.17s	333.178%	275.98%	17.17%
900	53.04s	334.30%	272.61%	18.4%
1000	58.91s	334.19%	266.71%	20.19%

This case study was evaluated with a test harness that creates a number of users with some initial deposit, that perform a number of transactions during their lifetime, after which they are deleted. As a range we chose to look at between 500 to 1000 users, with increments of 100. We measured the time taken for the test scenarios to execute:

(i) without monitoring;

(ii) with monitoring the original properties; and

(iii) with monitoring the residual properties.

We then computed the percentage of time saved by the residual analysis by computing the difference between time taken with the original monitoring as a percentage of the latter time overheads.

Table. 1 shows the results of this benchmarking, showing the increase in load to the transaction system on the number of users increasing. However the results show the monitoring overheads to remain relatively stable even given the increase in load. The residual analysis reduced the monitoring overheads significantly in this case (around 20%), however the remaining overheads remain large. We do not measure the time taken for CLARVA to produce the residuals since this was negligible.

We conclude that residual analysis can be useful in reducing the overheads associated with runtime verification. This is however conditional on what the residual analysis manages to prove and prune. This experiment's artifacts, along with other case studies, can be found in the CLARVA repository (Azzopardi, 2019).

5 RELATED WORK

CLARVA was inspired by CLARA (Bodden et al., 2010), *CompiLe-time Approximation of Runtime Analyses*, a tool for reducing the runtime overheads required for monitoring properties representable as a finite-state automaton. CLARA is a framework for static analyses that optimise the program event instrumentation required to monitor a property by analysing the control-flow of the program. The first two analyses included in CLARVA, see Figure. 2, in fact are inspired by existing CLARA analyses, while the third analysis is novel. Unfortunately CLARA is no longer maintained, while it uses an aspect weaver, abc (Allan et al., 2005), that is also no longer maintained. This motivated us to develop our own tool rather than extending CLARA with DATEs. Moreover CLARA only deals with residual instrumentation and does not analyse properties in any meaningful manner to encode the static guarantees given by its analyses, unlike CLARVA.

CLARVA currently only incorporates control-flow analysis for DATEs, since the program model (non-deterministic automata) does not maintain any knowledge about the program data-flow. Work already exists to analyse the data-oriented aspects of a DATE (i.e. transition guards) against the program, namely that of STARVOORS (Chimento et al., 2015) (*unified STAtic and Runtime Verification of Object-Oriented Software*) (Chimento et al., 2015). This tool uses a syntactic superset of DATEs, namely ppDATEs, that extend DATEs by allowing states to be tagged by Hoare triples over method calls. However, a ppDATE is semantically equivalent to a set of DATEs; in fact they are translated into DATEs for monitoring. STARVOORS uses the Java theorem prover KEY (Ahrendt et al., 2016) to prune these Hoare triples or to make the conditions weaker, reducing what must be proven at runtime by reducing the property. However STARVOORS does not consider attack program instrumentation reductions directly, unlike CLARVA.

There is other work that uses a similar notion of a property residual. (Dwyer and Purandare, 2008) analyses the control-flow of a program against a finite-state automaton to identify sequences of statements that always trigger the same behaviour in the property. A summary-based approach here is used where the instrumentation of such sequences is replaced by new event symbols representing entry and exit into it and appropriate transitions added to the automaton. This is different from our approach where we simply reduce the property and do not transform it.

Different work by (Lal et al., 2007; Beyer et al., 2018) creates a *residual program*, in effect the part of the program that was not proven safe by the analysis. In CLARVA we simply reduce instrumentation and do not change program behaviour, while in this other work the program is transformed into a smaller program.

6 CONCLUSIONS AND FUTURE WORK

In this paper we have introduced a tool, CLARVA, for the static analysis of monitors based on symbolic automata. This tool gives static guarantees by both reducing the monitor logic left to prove at runtime and by reducing the event instrumentation required of the program. CLARVA currently incorporates control-flow analyses of the program by analysis of properties against non-deterministic automata representations of program methods.

Future iterations of the tool are planned to extend the process to be able to deal with communication between different DATEs. This is envisioned to be possible through composing these communicating DATEs. Another area for improvement is the incorporation of data-flow analysis by extending the program model with data-oriented aspects, allowing for more precise analysis and smaller residual properties. We explored this theoretically in another publication (Azzopardi et al., 2020), however we have yet to implement this in CLARVA. We also plan to apply the same approach to different programming paradigms, including the LARVA variant for the Ethereum blockchain (Azzopardi et al., 2018).

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