Efficient and Precise Typestate Analysis by Determining Continuation-Equivalent States

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ABSTRACT

Typestate analysis determines whether a program violates a set of finite-state properties. Because the typestate-analysis problem is statically undecidable, researchers have proposed a hybrid approach that uses residual monitors to signal property violations at runtime.

We present an efficient novel static typestate analysis that is flow-sensitive, partially context-sensitive, and that generates residual runtime monitors. Our typestate specifications can refer to multiple interacting objects. To gain efficiency, our analysis uses precise, flow-sensitive information on an intra-procedural level only, and models the remainder of the program using a flow-insensitive pointer abstraction. Unlike previous flow-sensitive analyses, our analysis uses an additional backward analysis to partition states into equivalence classes. Code locations that transition between equivalent states are irrelevant and require no monitoring. This approach is simpler than previous approaches, nevertheless yields excellent precision and requires little analysis time.

We proved our analysis correct, implemented the analysis in the CLARA framework for typestate analysis, and applied it to the DaCapo benchmark suite. In half of the cases, our analysis determined exactly the property-violating program points. For another 25%, the analysis reduced the number of instrumentation points by large amounts, yielding significant speed-ups during runtime monitoring.

Categories and Subject Descriptors

D.2.4 [**Software Engineering**]: Software/Program Verification—*Validation*

General Terms

Algorithms, Experimentation, Performance, Verification

Keywords

typestate analysis, static analysis, runtime monitoring

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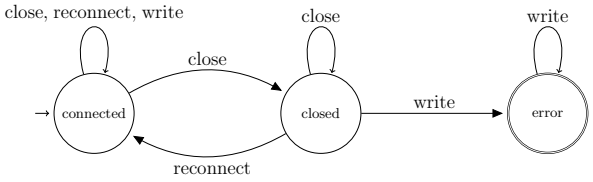


Figure 1: Finite-state machine for "Connection" property

1. INTRODUCTION

A typestate property [21] describes which operations are available on an object or even a group of inter-related objects, depending on this object's or group's internal state, the typestate. For instance, programmers must not write to a connection handle that is currently in its "closed" state. Figure 1 shows a non-deterministic finite-state machine for this property. It monitors a connection's "close", "reconnect" and "write" events and signals an error at its accepting state.

Typestate properties aid program understanding, and one can even define type systems [5, 14] that prevent programmers from causing typestate errors, or derive static typestate analyses [16] that try to determine whether a given program violates typestate properties. Unfortunately, the typestate-analysis problem is generally undecidable. Researchers have therefore proposed a hybrid approach [9, 10, 15] that uses static-analysis results to generate a residual runtime monitor. This monitor captures actual property violations as they occur, but only updates its internal state at relevant statements, as determined through static analysis.

A correct runtime monitor must observe events like "close" and "write" that cause a possible property violation, but also events like "reconnect" that may prevent the violation from occurring. Missing the former causes false negatives while missing the latter causes false positives, i.e., false warnings. Both is unacceptable, as runtime monitors must handle property violations exactly when they occur. A correct static analysis must therefore determine program locations that can trigger either kind of such "relevant" events.

In this work we present an efficient novel static typestate-analysis algorithm called Nop-shadows Analysis[1] that uses a forward and a backward pass to identify provably irrelevant code locations. For every program statement s of interest, the forward analysis determines the possible typestates that can reach s. The additional backward analysis partitions these states into equivalence classes. A program location

[1]The aspect-oriented-programming community uses the term "shadow" [17] to refer to instrumentation points.

that can only transition between equivalent states is irrelevant. This eases the burden on the programmer, as the programmer does not need to consider such irrelevant locations during manual code inspections. Moreover, we eliminate the monitoring instrumentation from these program locations, speeding up the residual runtime monitor. Our novel analysis is not only simpler than earlier approaches, it is also surprisingly precise, yet efficient.

Any precise typestate analysis has to be flow-sensitive and requires must-alias information: to determine that the code "c1.reconnect(); c2=c1; c2.write();" correctly uses the connection that c1 and c2 refer to, the analysis needs to know that c1 and c2 must point to the same object, i.e. that c1 and c2 must-alias. Such information is expensive to compute. To gain efficiency, our analysis computes flow information and must-alias information on an intra-procedural level only, and models the remaining program using a carefully designed flow-insensitive pointer abstraction.

To evaluate our approach, we have implemented our analysis in the CLARA framework for typestate analysis [7] and applied the analysis to the DaCapo benchmark suite [6]. Our results show that our lightweight abstractions are precise enough to exactly tell apart property-violating program points from irrelevant program points in half of the cases. For these cases, the analysis determines exactly the property-violating program locations. In another 25% of the cases, the analysis often identifies and disables large amounts of irrelevant program points. This eases manual code inspection and, as we show, speeds up the resulting residual monitor significantly. Our modest abstractions restrict the analysis time to a few minutes in most cases.

We proved our analysis correct. As we found out during this process, two analyses that we [10] and others [19] published previously are unsound. They fail to identify certain program locations that trigger violation-preventing events like "reconnect" above. As a consequence, the resulting runtime monitors may cause false warnings at runtime. This unsoundness is caused by the fact that traditional typestate analyses use a forward-analysis pass only and have no notion of equivalent states like our novel analysis does.

To summarize, this paper presents the following original contributions:

- a novel flow-sensitive static typestate analysis, called "Nop-shadows Analysis", that detects equivalent typestates to determine all statements that are relevant to causing or preventing a property violation,
- an implementation of the Nop-shadows Analysis in the CLARA framework that generates efficient residual runtime monitors,
- a set of experiments that shows that the analysis is both precise and efficient, and
- an explanation of why some other static analyses that use a forward-analysis pass only, generate unsound runtime monitors.

We structured the remainder of this paper as follows. We start off by giving a brief overview of the CLARA framework, in which we implemented our analysis. In Section 3 we give an example to illustrate the Nop-shadows Analysis itself. We explain the full analysis in Section 4, followed by our experiments in Section 5. In Section 6 we discuss how we improve over the state of the art. We conclude in Section 7.

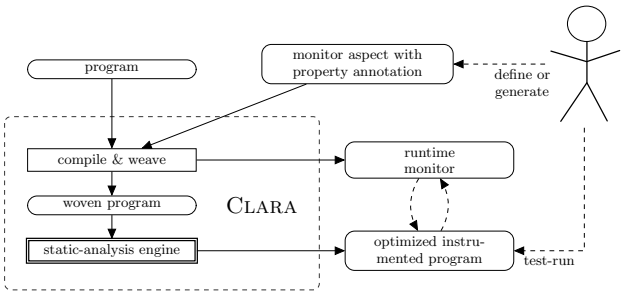


Figure 2: Overview of CLARA

2. THE CLARA FRAMEWORK

CLARA (CompiLe-time Approximation of Runtime Analyses) is a novel research framework for the implementation of hybrid typestate analyses. We developed CLARA to support easy implementation of the analysis that we present in this paper, among others. CLARA's major design goal is to de-couple the code-generation for efficient runtime monitors from the static analyses that convert these monitors into faster, residual monitors. In this work, we can only give an overview of CLARA. The author's dissertation [7] gives a more detailed account. CLARA is available at:

`http://bodden.de/clara/`

Figure 2 gives an overview of CLARA. With CLARA, the researcher first defines a set of typestate properties, denoting each property as an annotation to an AspectJ [3] aspect that implements a runtime monitor for the same property. Annotations directly encode a non-deterministic finite-state machine, just as the one in Figure 1. In the author's dissertation [7] we show how researchers can even use runtime-monitoring approaches like tracematches [2], JavaMOP [13], and others, to generate these annotated aspects automatically from high-level monitor specifications.

CLARA weaves the monitoring aspect into the program under test and emits helper classes that implement the runtime monitor as defined by the aspect. CLARA extends the AspectBench Compiler [4] for this purpose. CLARA then invokes its static-typestate-analysis engine. Researchers can add a number of static analyses to CLARA and have them applied in any order. These analyses obtain, through the finite-state machine defined in the aspect's annotation, enough information about the typestate property to precisely approximate the set of relevant instrumentation points. When an analysis determines that an instrumentation point is irrelevant to a property, i.e., the program can neither violate the property nor prevent a property violation at this point, then CLARA automatically disables the instrumentation for this property at this point. The result is an optimized instrumented program that updates the runtime monitor only at program points at which instrumentation remains enabled. The approach that we present in the following instantiates CLARA's static-analysis engine with a combination of two previously published supporting analyses and our novel Nop-shadows Analysis.

3. ANALYSIS BY EXAMPLE

We motivate our analysis using our running example: it is an error to write to a connection object that was closed, unless the connection was re-connected in between. Figure 1 shows the appropriate non-deterministic finite-state

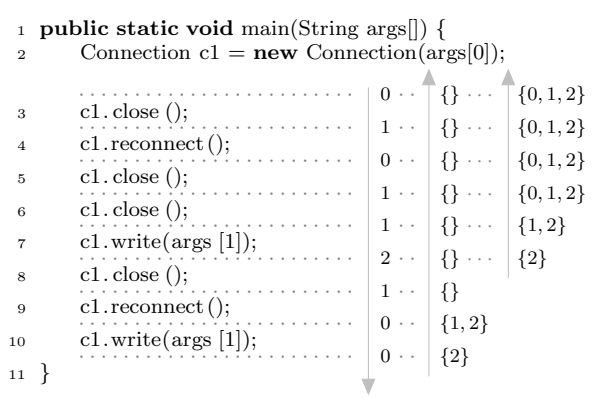
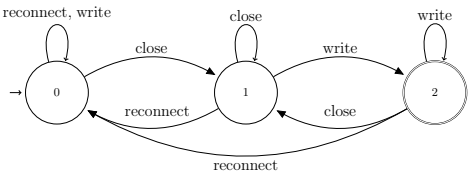
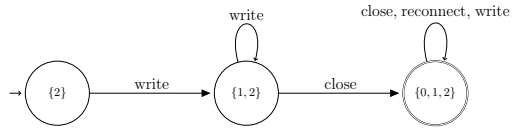


Figure 3: Simple example program using a single connection



(a) Deterministic finite-state machine $\mathcal{M}_{forward}$ for $\mathcal{L}$



(b) Deterministic finite-state machine $\mathcal{M}_{backward}$ for $\overline{\mathcal{L}}$

Figure 4: Finite-state machines for Connection example

machine that Clara obtains from parsing the annotation in the specified aspect. The state machine accepts events of type "close", "reconnect" and "write". When a "write" follows a "close", then the state machine moves into its error state. Clara represents error states as accepting states. Let us call the language that this state machine accepts $\mathcal{L}$.

Figure 3 shows a simple example program that uses a single connection, along with the analysis information that we compute (explanation follows). To keep the example simple, this program contains only straight-line code, no outgoing method calls that may change the connection's type state, and no aliasing. Our implementation, however, handles complete Java programs, including method calls, recursion, loops, exceptions and aliasing (see Section 4).

Our example program violates the connection property by closing the connection (even twice, at lines 5 and 6), and then writing to the connection (line 7). Note, though, that all other statements in this program are irrelevant to the property violation. In particular, one does not need to monitor the "close" and "reconnect" operations at lines 3 and 4 because they precede the violating fragment of the run. Conversely, the operations at lines 8 to 10 follow this fragment, and hence do not need to be monitored either. A little more subtle, even of the the two "close" events at lines 5 and 6, it is correct to omit monitoring one of them. (But not both!) The static analysis that we present in the following will eliminate the instrumentation at exactly those shadows that we just identified as irrelevant—"nop shadows", as we call them. Instrumentation will only remain in lines 5 and 7, or 6 and 7—an optimal result for this program.

Example application of analysis algorithm.

For a method containing n shadows, the Nop-shadows Analysis consists of up to n "rounds", where each round identifies a single nop shadow until no further nop shadows can be identified. Each round consists of a forward and a backward pass. The forward pass computes for every statement s the typestates that can reach s. The backwards pass conversely computes classes of states from which the property state machine can reach a violating state using the remainder of the program execution that follows s.

The forward pass uses a determinized version of the finite-state machine from the property specification. Figure 4a shows this state machine for our example. In the following, we will call this state machine $\mathcal{M}_{forward}$. We number $\mathcal{M}_{forward}$'s states for presentation purposes. The forward analysis starts off in this state machine's initial state 0 and then updates the state according to the shadows that it encounters during analysis. In Figure 3, we show the states that the forward analysis computes before and after each statement, next to the downward arrow. For instance, the "close" statement at line 3 changes the typestate from 0 to 1. Importantly, at property violations, e.g. at line 7, the analysis will reach the violating state 2.

The backward pass, on the other hand, uses a deterministic finite-state machine for $\mathcal{L}$'s mirror language $\overline{\mathcal{L}}$. For any word $w = w_1 \dots w_n \in \mathcal{L}$, we define the mirror word $\overline{w}$ as $\overline{w} := w_n \dots w_1$. The mirror language $\overline{\mathcal{L}}$ is defined by $\overline{\mathcal{L}} := \{\overline{w} \mid w \in \mathcal{L}\}$. While it would be sound for our backwards pass to use any finite-state machine that accepts the language $\overline{\mathcal{L}}$, we specifically use the state machine that we obtain by (1) inverting $\mathcal{M}_{forward}$ (by flipping all edges and swapping initial and accepting states), and then (2) determinizing this finite-state machine again. The resulting state machine is minimal for $\overline{\mathcal{L}}$ [12]. As we will explain in Section 4, a minimal finite-state machine yields a more precise analysis result because in this state machine equivalent states are collapsed into a single state. For our example, Figure 4b shows the finite-state machine that we obtain this way. (We omit the "sink" state that represents the empty state set.) We call this automaton $\mathcal{M}_{backward}$. Note that we labeled each of this automaton's states with the set of states of $\mathcal{M}_{forward}$ that this state represents. These labels are important: our analysis will compare states from the forward analysis with states from the backward analysis, and labeling $\mathcal{M}_{backward}$'s states with their equivalent states of $\mathcal{M}_{forward}$ eases this comparison.

According to the semantics of Clara's state-machine notation, a single program run can cause multiple property violations. Each violating sub path of the execution path starts at one of the program's possible entry points and ends at what we call a "final shadow". A final shadow is a shadow that is labeled with a symbol that leads into an error state, like "write" in our example. Therefore, in the example, we apply the backwards analysis two times, starting at both write statements (lines 10 and 7). In Figure 3, we show the analysis result for both backwards-analysis runs on the right-hand side. For instance, the close statement at line 6 changes the typestate from the state labeled with $\{1, 2\}$ to the state labeled with $\{0, 1, 2\}$. The same statement also causes the second analysis configuration, holding the sink state $\{\}$, to loop.

Nop-shadow condition. We now explain how we combine the forward and backward-analysis information to identify nop shadows. Let *source*(s) be the state that the forward analysis computed just before a statement s, *target*(s) the state for the location just after s, and *futures*(s) the set of state sets that the backwards analysis computed for just after s. For instance, for the close statement at line 5 of Figure 3 we have:

$$\begin{aligned} \textit{source}(\text{line 5}) &= 0 \\ \textit{target}(\text{line 5}) &= 1 \\ \textit{futures}(\text{line 5}) &= \{\ \{\}, \{0,1,2\}\ \} \end{aligned}$$

Because we compute *futures*(s) using a deterministic finite-state machine for $\overline{\mathcal{L}}$, the sets in *futures*(s) represent equivalence classes. For instance, the set $\{0,1,2\}$ represents the fact that, using the remainder of the program execution, one will reach a property violation from $\mathcal{M}_{forward}$'s states 0, 1 or 2 *either way*. By using a minimal state machine we assure optimality: when two states q_1 and q_2 are equivalent, then *futures*(s) will contain a state set Q with $\{q_1, q_2\} \subseteq Q$.

In the following, for two states q_1 and q_2 we say that q_1 and q_2 are equivalent and write $q_1 \equiv q_2$ if the following holds:

$$\forall Q \in \textit{futures}(s).\ q_1 \in Q \leftrightarrow q_2 \in Q$$

A shadow is a nop shadow when it transitions between states in the same equivalence class. Let us denote by F the set of accepting, i.e., violating states of $\mathcal{M}_{forward}$. Then we call a shadow at a statement s a "nop shadow" if:

1. *source*(s) $\equiv$ *target*(s), and
2. *target*(s) $\notin F$.

The first case states that the shadow transitions between states that are in the same equivalence class. Hence, monitoring the shadow appears unnecessary. (Note that, as a special case, this condition handles looping: when *source*(s) = *target*(s) then Condition 1 holds trivially.) However, there is one exception that we need to consider, and which we handle in Condition 2: When *target*(s) $\in F$, then the shadow directly triggers the runtime monitor. According to CLARA's monitoring semantics, a monitor must signal repeated property violations every time the violation occurs. (This is useful when the monitor executes error-handling code.) For instance, on "c.close();c.write();c.write()" the monitor should signal a violation after both "write" events. However, the second "write" event does not change the typestate; we have *source*(s) = *target*(s) = 2. Therefore, Condition 1 holds although the statement is not a nop shadow. Adding Condition 2 handles this corner case.

Need for re-iteration. The diligent reader may wish to verify that this condition indeed identifies all the statements of our example (Figure 3) as nop shadows, except for the write statement at line 7. This means that it is correct to not observe any *single* one of the shadows at these statements. However, it would be incorrect to state that, for instance, *both* the shadows at lines 5 and 6 are irrelevant: one needs to observe one of these two shadows; otherwise the runtime monitor will not reach its violating state at line 7. The problem is that after identifying any particular single shadow as a nop shadow, by disabling this shadow, we change the program's transition structure. Therefore, after identifying any shadow at any statement s as a nop shadow, we recompute both the forward and the backward-analysis information, this time *ignoring* the shadow at s. This yields updated analysis information that is again sound to use for identifying further nop shadows. We re-iterate until we reach a fixed point, i.e., cannot identify any nop shadow any more. The number of iterations is bounded by the number of shadows that the current method holds. In our example, the 7th iteration will expose no further nop shadows, and instrumentation will only remain in lines 5 and 7, or 6 and 7—an optimal result for this program.

4. NOP-SHADOWS ANALYSIS

We next explain how we handle the general analysis problem, involving loops, outgoing method calls, recursion, exceptions and aliasing.

4.1 Supporting analyses

We instantiated CLARA's static-analysis engine with a sequence of three analysis stages where each stage is strictly more precise but also more expensive to compute than preceding stages. Each stage can identify and disable nop shadows and can therefore speed up later stages that do not need to consider these shadows any more. The stages that we use are (1) the syntactic Quick Check, (2) the flow-insensitive Orphan-shadows Analysis, and (3) the novel flow-sensitive Nop-shadows Analysis that we present in this paper.

The Quick Check [8] is very fast because it uses syntactic information only. It rules out instrumentation points if certain automaton symbols match *nowhere* in the program. For instance, in our Connection example, if the program never writes to a connection then the program cannot violate the property and the Quick Check can disable all instrumentation. CLARA then bypasses the other stages.

The Orphan-shadows Analysis [8] performs a similar analysis separately for each abstract object (or combination of such objects), where we model abstract objects through object representatives [11][2]. For instance, consider a program that connects and closes connections but never writes to some of those. Because the program cannot violate the connection property at places that provably only refer to such "read-only" connections, the Orphan-shadows Analysis removes instrumentation from these program locations.

These first two analysis stages disable shadows that are obviously irrelevant. CLARA applies the Nop-shadows Analysis to all interesting cases, i.e., the ones in which shadows remain enabled after applying these first two stages. The Nop-shadows Analysis is strictly more precise than stages (1) and (2) because it uses flow-sensitivity and must-alias information. Both are necessary for precise typestate analysis but also expensive to compute. We therefore designed the Nop-shadows Analysis to compute must-alias information and flow information on an intra-procedural level only. One may think that such an analysis would have to be quite imprecise, but before we designed our analysis, we manually investigated the instrumentation points that remained active after the flow-insensitive Orphan-shadows Analysis had already been applied to our benchmarks. We found that, in most cases, when paired with coarse-grain interprocedural summary information that the Orphan-shadows

[2]Object representatives combine flow-insensitive whole-program points-to sets with intra-procedural flow-sensitive alias information. We compute points-to sets with Sridharan and Bodík's context-sensitive points-to analysis [20].

Analysis had already computed anyway, intra-procedural analysis information was sufficient to determine irrelevant instrumentation points. The results that we present in Section 5 confirm these findings.

4.2 Actual Nop-shadows analysis

The Nop-shadows Analysis identifies nop shadows on a per-method, per-property basis. For each statement s in the current method, the forward analysis computes the possible states $source(s)$ and $target(s)$. The backward analysis, on the other hand, computes $futures(s)$. The forward and backward analysis are virtually dual instances of a general worklist algorithm that propagates "configurations" through the method's control-flow graph. In our implementation, a configuration $c = (Q, b)$ combines an automaton-state set Q with a variable binding b. When c is associated with a statement s, then the state set Q holds all possible typestates just before executing s. The binding b describes the object(s) to which this state set belongs. In our connection example, for any given statement s there could exist multiple connections that are in different typestates when s executes. The variable bindings help distinguish these different typestates. A variable binding maps one or more variables from the typestate specification to object representatives that model the runtime objects that these variables are bound to. The treatment of variable bindings is quite intricate, but other aspects of our analysis are more interesting. Hence, we decided to ease our presentation by abstracting from variable bindings and instead assuming that we perform typestate analysis only for one single object representative. For the remainder of this paper we therefore assume that a configuration is just a set of automaton states, without any binding. The author's dissertation [7] gives a complete treatment including variable bindings, with proof.

The Nop-shadows Analysis propagates configurations for both the forward and backward analysis using a general worklist algorithm that we show as Algorithm 1. In this algorithm, the syntax $f[x \mapsto y]$ denotes the function that is equal to f on all values v, except for x, in which case it returns y:

$$f[x \mapsto y] := \lambda v \begin{cases} y & \text{if } v = x \\ f(v) & \text{otherwise} \end{cases}$$

The algorithm first initializes a worklist wl, which is essentially a set of "jobs", where each job is an entry $(stmt, cs)$, with $stmt$ a statement and cs is a set of configurations, i.e., a set of set of automaton states. The algorithm further initializes two mappings *before* and *after* that store the configurations that have been computed so far before, respectively after each statement. These sets allow us to perform a terminating fixed point iteration.

Next the algorithm iterates through its worklist. For every job $(stmt, cs)$, the algorithm first updates $stmt$'s before-set. Then, when a statement holds no shadow, we just leave the configurations unchanged (line 6 in Algorithm 1). Otherwise, we compute (line 7), for every new configuration $c \in cs$ and $shadow$ at $stmt$, successor configurations using the supplied transition function δ. To compute the transition, the algorithm accesses the shadows's unique label $label(shadow)$. (In our running example, this label could be "close", "reconnect" or "write".) To allow the analysis to later-on compare state labels of $\mathcal{M}_{forward}$ with the state-set

Algorithm 1 $worklist(initial, succ_{cfg}, succ_{ext}, \delta)$

```
1:  wl := initial
2:  before := after := λstmt. ∅
3:  while wl non-empty do
4:    pop job (stmt, cs) from wl
5:    before := before[stmt ↦ before(stmt) ∪ cs]
6:    cs' := { cs   if shadows(stmt) = ∅
             { ∅    otherwise
7:    for c ∈ cs,  shadow ∈ shadows(stmt) do
8:      c' := ⋃_{q∈c} { δ(q, label(shadow)) }
9:      cs' := cs' ∪ {c'}
10:   end for
11:   cs_new := cs' − after(stmt)
12:   if cs_new non-empty then
13:     after := after[stmt ↦ after(stmt) ∪ cs_new]
14:     for stmt' ∈ succ_cfg(stmt) do
15:       wl := wl ∪ {(stmt', cs_new)}
16:     end for
17:     for stmt' ∈ succ_ext(stmt) do
18:       wl := wl[stmt' ↦ wl(stmt') ∪
                reaching(cs_new, relatedShadows(stmt))]
19:     end for
20:   end if
21: end while
```

labels that $\mathcal{M}_{backward}$ uses, we determinize state machines on-the-fly: line 8 computes the unique set of successor states.

The algorithm then updates $stmt$'s after set and associates new jobs with two different kinds of successor statements. First, in lines 14–16, the algorithm adds new jobs containing the successor configurations cs_{new} for any statement that is a successor of $stmt$ in m's control-flow graph (as determined by $succ_{cfg}$). Lines 17–19 handle inter-procedural control flow through an "external-successor function" $succ_{ext}$. After all, the typestate of objects that the current method m refers to may also be changed by methods other than m.

4.2.1 The external-successor function

Figure 5 visualizes both successor functions. We show the current method m as a box. The method contains two invoke statements. The first statement resembles a potentially-recursive call, the second one a provably non-recursive call. The dashed arrows denote the successor function $succ_{cfg}$, which is given by m's control-flow graph. In addition, the solid arrows show the second, inter-procedural-successor function, $succ_{ext}$. During the execution of m, invoke expressions within m may cause methods to be called. These calls either can or cannot transitively perform a recursive call back into m. When the call may be recursive, then configurations that we computed for this call site can reach m's entry statement, see arrow (1). Conversely, for configurations that we computed for any of m's exit statements, we need to propagate these configurations back to any potentially-recursive call site within m, see (2). At compile time, we can usually only determine that a method call may be recursive, not that it must be. Hence, we also need to propagate configurations from the call site to after itself, see (3a). For calls that are provably not recursive (as determined by a call graph), it suffices to propagate configurations past the call site itself, see (3b). Lastly, we need to

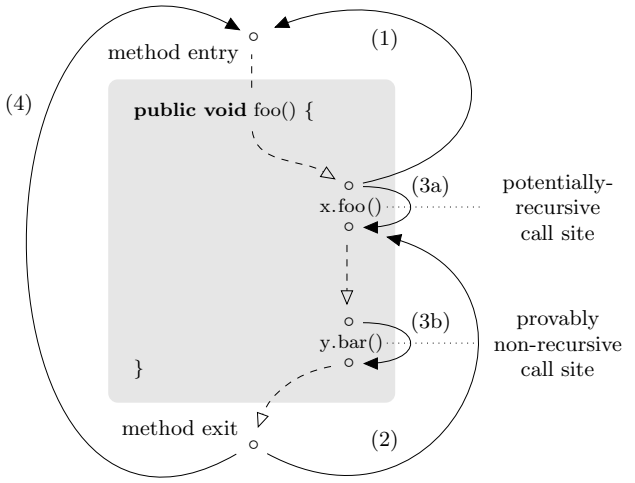


Figure 5: Functions $succ_{cfg}$ (dashed) & $succ_{ext}$ (solid)

take into account the case in which method m returns (either normally or by throwing an exception—$succ_{cfg}$ handles both cases), and then re-executes. To model this case we propagate configurations from m's exit(s) to its entry, (4).

In line with Figure 5, we define $succ_{ext}$ as follows. Let $heads(m)$ be the set of entry statements of m, and $tails(m)$ the set of m's exit statements[3]. Further, let $recCall(m)$ be the set of potentially-recursive invoke statements of m, and $nonRecCall(m)$ the set of provably non-recursive invoke statements respectively. Then:

$succ_{ext} :=$

$$\lambda s.\ \begin{cases} succ_{cfg}(s) & \text{if } s \in nonRecCall(m) \\ heads(m) \cup succ_{cfg}(s) & \text{if } s \in recCall(m) \\ succ_{cfg}(recCall(m)) \cup heads(m) & \text{if } s \in tails(m) \\ \emptyset & \text{otherwise} \end{cases}$$

When propagating configurations along a $succ_{ext}$-edge, it is not correct to just copy the configurations from the edge's source to it's target. Note that between any two executions of m, other methods may execute and cause transitions in the monitoring state machine. To soundly model these potential transitions by "other methods", Algorithm 1 associates in line 18 with any inter-procedural successor not just the set of new configurations cs_{new} but instead the set of configurations $reaching(cs_{new}, relatedShadows(stmt))$. *reaching* computes successor configurations in a very efficient, flow-insensitive way. Let cs be a set of configurations, ss a set of shadows and $l := \{label(s) \mid s \in ss\}$. We define $reaching(cs, ss)$ as the smallest set for which holds that:

- $cs \subseteq reaching(cs, ss)$, and
- $\forall c \in reaching(cs, ss)\ \forall a \in l :$
  $\{\delta(q, a) \mid q \in c\} \in reaching(cs, stmt)$.

Hence, *reaching* computes the configurations that one can reach from cs by executing zero or more shadows from ss.

The function $relatedShadows(stmt)$, that the Algorithm further uses in line 18, computes the set of all shadows related to $stmt$. We define this set as follows. For any invoke statement $stmt$ (potentially recursive or not), the set $relatedShadows(stmt)$ contains all shadows in all methods transitively reachable through $stmt$, except for the ones in m itself. After all, these are all the shadows that one can reach before reaching m's entry statement again. Otherwise, i.e., if $stmt$ is a head or tail of m, then $relatedShadows(stmt)$ contains all shadows in the program, except for the ones in m. (Our implementation further narrows down related shadows by comparing each shadow's variable binding to the binding stored in the configuration.)

It is worthwhile noting that, because we compute the expression $reaching(cs_{new}, relatedShadows(stmt))$ for each statement separately, we gain a certain amount of context-sensitivity. While shadows inside a certain method m' (with $m' \neq m$) may be relevant to one statement of m they may be irrelevant to other statements in m, and by recomputing the above function we properly distinguishes such cases.

4.2.2 Initializing the worklist algorithm

We next explain how we initialize the worklist algorithm. The initialization depends on whether we perform a forward or backwards analysis. In forward-analysis mode, $succ_{cfg}$ is simply the successor function of m's control-flow graph, and $succ_{ext}$ is the inter-procedural successor function as defined above; the function δ is the transition function of $\mathcal{M}_{forward}$. For the backwards analysis we simply reverse both successor functions and the transition function, flipping their edges.

We determine the set *initial*, which Algorithm 1 uses to initialize its worklist in line 1, as follows. Let q_0 be the initial state of $\mathcal{M}_{forward}$. Assume we analyze method m, and let *otherShadows* be the set of shadows outside of m. For the forward analysis, we define *initial* as:

$$\{\ (h, reaching(\{q_0\}, relatedShadows(h))) \mid h \in heads(m)\ \}$$

This set of initial jobs associates with m's entry statement all configurations that are reachable from the initial configuration $\{q_0\}$ by executing every possible method, except m itself. (Note that our inter-procedural-successor function $succ_{ext}$ already handles the case where m re-executes.)

We generate initial jobs for the backwards analysis in a similar but not identical way. A violating trace can only start at the beginning of the program, but it can end (causing a violation) in the current method m itself, or in another method (either with m on the call stack or not). We generate initial jobs to cover these three cases. Due to space limitations we give a formal definition in the accompanying dissertation [7, Section 5.2.3.4].

4.2.3 Removing nop shadows

The analysis information directly provides us with *source*, *target* and *futures* for every statement. We use this information to identify and disable a nop shadow if possible, and then re-iterate until we can find no further nop-shadows for this method. In our benchmark set we had to iterate ten times or less for all but four methods. When we reach the fixed point, we proceed with the next method. When all methods are processed, we apply the flow-insensitive Orphan-shadows Analysis again and then re-iterate the whole Nop-shadows Analysis. This is because disabling a shadow in one method may render shadows in other methods irrelevant. It seems to be always sufficient to iterate this outer loop two to three times. When this loop reaches a fixed point we stop.

[3]Usually, $heads(m)$ will be a singleton set but because our backwards analysis operates on a reversed control-flow graph where heads become tails and tails become heads, $heads(m)$ can contain more than one element in this setting.

5. EXPERIMENTS

To validate our approach, we verified a set of twelve typestate properties over ten benchmark programs of the DaCapo benchmark suite [6]. This lead to 120 property/benchmark combinations. 43 of these combinations (36%) were "interesting" to us in the sense that instrumentation remained after applying the first two, previously published, analysis stages. The 43 combinations comprised eight out of the original twelve properties. Table 1 explains these properties. We applied the Nop-shadows Analysis to these 43 combinations.

Table 2 summarizes our analysis results. The table reports the fraction of "final shadows" that the analysis identified as nop shadows. A shadow is final if it can complete a property violation. For instance, in the Connection example, the final shadows are exactly all "write" shadows. The number of final shadows thus corresponds to the number of program points that may trigger property violations. The fraction of shadows that our analysis identified as nop shadows appears in white. In gray we show the fraction of shadows which are known to trigger actual violations at runtime. The remaining black slice represent shadows that remain active even after analysis, either due to analysis imprecision or due to actual property violations.

For 18 out of these 43 combinations (41%), our novel Nop-shadows Analysis was able to identify all shadows as irrelevant and therefore proved that the program cannot violate the stated property. These cases appear as all white circles. In four other cases, shadows remained enabled, but only because they do trigger a property violation. These cases appear as gray circles. In other words, the analysis gave exactly the correct result, with no false positives, in half of the cases. In eleven cases, the analysis was unsuccessful and did not identify any nop shadow (black circles).

In the remaining ten cases, the analysis removed a sometimes significant amount of shadows. This may speed up runtime monitoring for these cases, depending on whether the test run exercises these shadows a lot. For our experiments we used monitoring aspects generated from tracematches [2]. Table 2 gives qualitative information about the residual monitor's runtime overhead through the ring that surround each circle. (The dissertation [7] gives the full data.) Interestingly, the number of remaining shadows does not necessarily correspond directly to the resulting runtime overhead. For instance, a single shadow remains in antlr-Writer, but this one shadow executes so often that it causes a runtime overhead of more than 15%. chart-FailSafeIterMap, on the other hand, contains 38 residual shadows (no improvement), but there is no observable overhead. This is an ideal candidate for residual runtime monitoring. Altogether, after applying the Nop-shadows Analysis, only nine combinations remain that have a significantly perceivable overhead of more than 15%. Most combinations show zero overhead, five combinations show an overhead of below 15%, which we think is negligible at least at development time.

Our analysis works well on the antlr, fop, hsqldb, luindex, lusearch and xalan benchmarks. Most of the potential false positives (black in the figure) appear only because the benchmarks use reflection. Due to a known deficiency [1], Java's `Cloneable` interface contains no public declaration of a `clone()` method. Therefore, Java's type system may prevent clients from calling `clone()` even on `Cloneable` objects. chart uses reflection to call the `clone()` method on objects that implement the `Cloneable` interface. Because chart clones collections, our points-to analysis has to safely assume that the collections could be of any type, including `EmptySet`, which, as a singleton object, is stored in a static field, causing our analysis to loose all context information. bloat, jython and pmd cause similar problems.

FailSafeEnum	do not update a vector while iterating over it
FailSafeEnumHT	do not update a hash table while iterating over its elements or keys
FailSafeIter	do not update a collection while iterating over it
FailSafeIterMap	do not update a map while iterating over its keys or values
HasNextElem	always call hasMoreElements before nextElement on an Enumeration
HasNext	always call hasNext before calling next on an Iterator
Reader	do not use a Reader after its InputStream was closed
Writer	do not use a Writer after its OutputStream was closed

Table 1: Relevant typestate properties and their names

There appear to be only few cases where our analysis is too imprecise because of its design. For example, two actually irrelevant final shadows remain enabled in hsqldb with Reader and Writer. These false positives occur because xalan uses different methods to open, close and write to streams. A fully inter-procedural analysis could rule out possible violations in these cases. However, we found that, due to its intra-procedural nature, the Nop-shadows Analysis has an interesting property: the analysis revealed missing pre-conditions on xalan's methods. For instance, the write-calling method is missing the pre-condition that the argument file should not be in state "closed". In future work we plan to use this information to support program understanding and to further enhance precision.

Detected property violations. When manually inspecting the remaining shadows, we found several actual property violations. bloat violates Writer because it contains a method that writes to a file handle that it then closes. When called multiple times, this method will violate the property. We could not confirm whether certain runs of bloat may actually call this method multiple times. jython sometimes violates Reader by closing a stream prematurely. jython then catches the resulting exception and returns null. xalan violates HasNextElem by calling `nextElement` without a preceding call to `hasMoreElements`. Nevertheless, the program is safe because it checks the size of the underlying vector to assure that the calls are legal. The property specification is too simplistic in this setting. Several benchmarks violate FailSafeEnum and FailSafeEnumHT. These benchmarks do indeed modify vectors (or hash tables) while iterating over them. This can lead to unexpected behavior. With iterators this does not usually happen because iterators have fast-fail semantics and will throw a `ConcurrentModificationException` in such situations. luindex violates the FailSafeIter pattern. This error probably remained undetected because it only occurs on quite unlikely execution paths. The author's dissertation [7] provides more details.

Analysis time. Our analysis time is clearly dominated by the time it takes to compute the supporting analyses that the Nop-shadows Analysis requires. Constructing a call graph and context-sensitive points-to sets took about two and a half minutes on average. The Nop-shadows Analysis

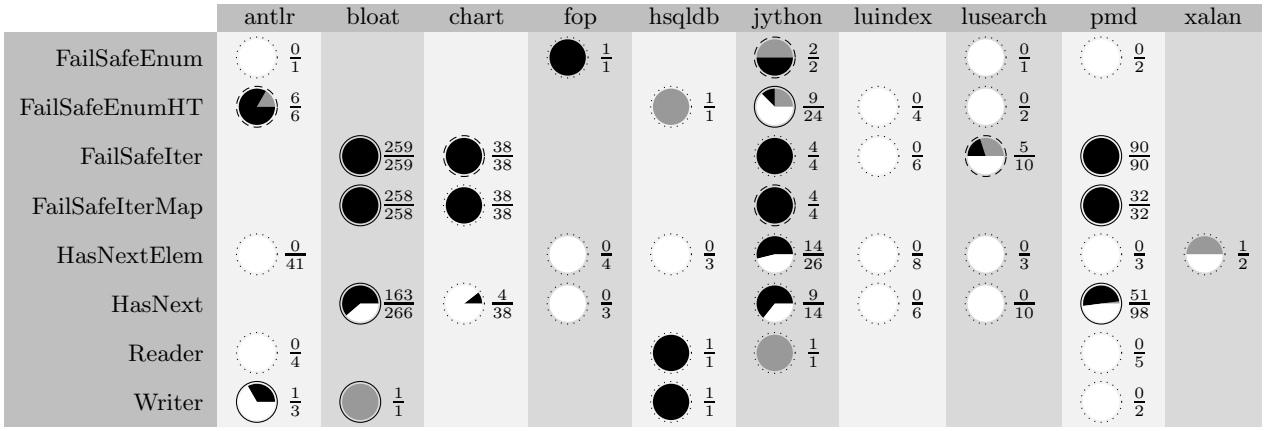


Table 2: Final shadows that may violate a property. White slices represent shadows that the Nop-shadows Analysis identified as irrelevant. Black slices represent shadows that we fail to identify as irrelevant, due to analysis imprecision or an actual violation. Gray slices represent actual property violations that we found through manual inspection. The outer rings represent the residual monitor's runtime overhead. Solid: overhead $\geq 15\%$, dashed: overhead $< 15\%$, dotted: no overhead.

itself took under 50 seconds on average. This time includes all re-iterations of the Orphan-shadows Analysis and Nop-shadows Analysis that CLARA performs. In 90% of the cases, the analysis finished in under one minute. By far the worst case was bloat-FailSafeIter, for which this analysis stage took 19 minutes. bloat is notoriously hard to analyze [8, 9, 19].

Limitations and threats to validity.

We identified the following limitations of our approach. All DaCapo benchmarks load classes using reflection. Static analyses like ours have to be aware of these classes so that they can construct a sound call graph. We wrote an AspectJ aspect that would print at every call to `forName` and a few other reflective calls the name of the class that this call loads and the location from which it is loaded. We further double-checked with Ondřej Lhoták, who compiled such lists of dynamic classes earlier. We then provided Soot [22] (which is part of CLARA) with this information. The resulting call graph is sound for the program runs that DaCapo performs. Obtaining a call graph that is sound for all runs may be challenging for programs that use reflection.

For eclipse we were unable to determine where dynamic classes are loaded from. eclipse loads classes not from JAR files but from "resource URLs", which eclipse resolves internally, usually to JAR files within other JAR files. Soot currently cannot load classes from such URLs and that is why we omit eclipse in our experiments. The jython benchmark generates code at runtime, which it then loads. We did not analyze this code and so made the unsafe assumption that this code would not cause any typestate changes.

Otherwise, the internal validity of our experiments is high because we directly measure the number of final shadows before and after the analysis. The final shadows are exactly the points that programmers would first inspect when checking possible property violations. Hence, reducing the number of final shadows will reduce the burden on the programmer. This is especially true when eliminating all final shadows, thus proving that the program cannot violate the property.

To measure the runtime overheads precisely, we extended the DaCapo harness with a custom driver class. With this driver class, DaCapo first executes a warm-up run and then re-runs the benchmark multiple times until the relative standard deviation of the determined runtimes drops below 3% (but at least 5 times and at most 20 times). Then we report the arithmetic mean of these runs. DaCapo's standard driver only measures a single benchmark run, which has caused misleading results for us in the past.

The external validity is limited by our choice of benchmarks. However, the DaCapo benchmarks are a realistic, representative set of medium-sized to large-scale applications. The suite contains both "well-behaved" benchmarks that are free of reflection and benchmarks that are harder to analyze due to reflection. Our analysis excels on the former, however, further work is required to handle the latter more effectively. We plan to address these problems by simulating reflection in a more fine-grained manner.

6. RELATED WORK

Strom and Yemini [21] were the first to suggest the concept of typestate analysis. In the last few years, researchers have presented several new approaches with varying cost/precision trade-offs. In the following we describe the approaches that are most relevant to our work. We distinguish type-system based approaches, static verification approaches and hybrid verification approaches.

Type-system based approaches.

Type-system based approaches define a type system and implement a type checker. This is to prevent programmers from compiling a potentially property-violating program in the first place and gives the advantage of strong static guarantees. On the other hand, the type checker may reject useful programs that statically appear to violate the stated property but will not actually violate the property at runtime. Our approach allows the programmer to define a program that may violate the given safety property. Our analysis then tries to verify that the program is correct, and when this verification fails it delays further checks until runtime.

Bierhoff and Aldrich [5] present an intra-procedural type-system based approach that enables the checking of typestate properties in the presence of aliasing. The author's approach aims at being modular, and therefore abstains from potentially expensive whole-program analyses like ours. To

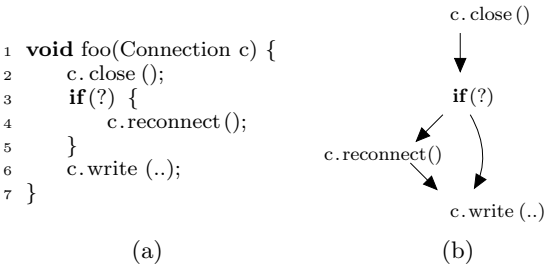


Figure 6: Example exposing unsoundness in earlier hybrid typestate analyses

be able to reason about aliases nevertheless, Bierhoff and Aldrich associate special access permissions with references. Access permissions allow the type checker to reason about a reference locally. The author's current approach assumes that a program contains information about access permissions and also typestate changes in the form of special program annotations. Our approach does not require any program annotations; it is fully automatic.

DeLine and Fähndrich's approach [14] is similar in flavor to Bierhoff and Aldrich's but uses a more restrictive abstraction of aliases that allows for less flexible calling conventions for typestate-changing methods. The authors implemented their approach in the Fugue tool for specifying and checking typestates in .NET-based programs. As in Bierhoff and Aldrich's approach, DeLine and Fähndrich assume that a programmer (or tool) has annotated the program under test with information about how calls to a method change the typestate of the objects that this method references.

Static analysis approaches.

In this section we describe approaches that, unlike type systems, perform a whole-program analysis and, unlike hybrid approaches, have no runtime component.

Fink et al. present a static analysis of typestate properties [16]. Their approach, like ours, uses a staged analysis which starts with a flow-insensitive pointer-based analysis, followed by flow-sensitive checkers. The authors' analyses allow only for specifications that reason about a single object at a time, while we allow for the analysis of multiple interacting objects. Fink et al.'s algorithms only determine "final shadows" that complete a property violation (like "write" in our example) but not shadows that initially contribute to a property violation (e.g. "close") or can prevent a property violation (e.g. "reconnect"). Therefore, these algorithms are unsuitable for generating residual runtime monitors.

Hybrid analysis approaches.

In own previous work [10] we presented a hybrid typestate analysis that was, like the Nop-shadows Analysis, also flow-sensitive on an intra-procedural level only, and used a flow-insensitive abstraction of the remainder of the program. However, unlike the Nop-shadows Analysis, the earlier analysis used a forward-analysis only. A forward analysis can only approximate the possible *source* and *target* states of a statement s but not the *futures*.

Consider the property-violating program in Figure 6a. Remember that a correct runtime monitor must not only observe events at property-violating shadows like the "close" shadow at line 2 and the "write" shadow at line 6, but also at shadows that may prevent a violation, like the "reconnect" shadow at line 4. A design flaw caused our earlier forward analysis to mistakenly disable certain violation-preventing shadows like the reconnect shadow in this example. To determine relevant shadows, this earlier analysis used what we called a "shadow history". The shadow history at a statement s is the set of all shadows on the control flow that reaches s. When determining that the program may reach an error state at s, the analysis would then commit the shadow history at s to a global set of "relevant shadows" that need to be monitored at runtime.

In Figure 6b we show the control-flow graph for the example program. As the graph shows, due to the if statement, the analysis will reach the write statement along two different branches. Along the left branch, the analysis determines that the connection is in its "connected" state when being written to, and therefore no shadow history is committed, as no violation can occur. Along the right branch, the analysis determines that the connection will be in state "closed" when reaching the write at line 6. Hence, the analysis will commit the shadow history. However, the shadow history along this branch contains the disconnect and write statements only, because the reconnect occurs on the other branch! The residual runtime monitor for this property will therefore miss any possible reconnect events, and may therefore signal false runtime warnings. We proved that our novel Nop-shadows Analysis causes neither false warnings nor missed violations.

Compared to the earlier analysis, the Nop-shadows Analysis uses an entirely different design, introducing the notion of equivalent states. This has several advantages. Firstly, by using a backward analysis, we obtain directly at every statement s exactly the necessary information that is necessary to decide whether s is relevant. Secondly, the resulting analysis is simpler, which greatly eased our soundness proofs. Further, the new analysis is more efficient: because we abstain from using a shadow history we obtain a smaller abstraction. The new abstraction only associates states with bindings, while the old analysis further associated this information with sets of shadows. Its smaller abstraction helps our new analysis to a faster-terminating fixed-point iteration.

Naeem and Lhoták present a fully context-sensitive and flow-sensitive inter-procedural whole-program analysis for typestate-like properties of multiple interacting objects [19]. The analysis that Naeem and Lhoták present can be seen as a generalized version of our own earlier analysis [10]. Naeem and Lhoták's analysis is fully inter-procedural. This can yield enhanced precision in cases where combinations of objects that are relevant to a given specification are used by multiple methods. Our benchmark set showed some instances where this additional information would have been helpful, but not many. It even holds that, although our analysis is intra-procedural only, there are some instances where our analysis is more precise than Naeem and Lhoták's. This is due to the highly context-sensitive points-to sets that we compute. It is important to note that Naeem and Lhoták also use shadow histories to determine relevant instrumentation points. Unfortunately, their analysis therefore suffers from the same unsoundness problem that we described above. Naeem and Lhoták had proven their entire analysis sound except for the use of shadow histories [18].

It is due to this unsoundness problem that we did not

compare our results to these of Naeem and Lhoták or to our own earlier results. The results would be incomparable: an unsound analysis could appear more effective by falsely ruling out shadows that are actually relevant. It would be interesting to use Naeem and Lhoták's abstractions to determine equivalent states like we do. This would be non-trivial because in our approach we re-compute all intra-procedural analysis information after disabling any single shadow. In an inter-procedural setting this would be too expensive.

Dwyer and Purandare use existing typestate analyses to specialize runtime monitors [15]. Their work identifies "safe regions" in the code using a static typestate analysis. Safe regions can be methods, single statements or compound statements (e.g. loops). A region is safe if its deterministic transition function does not drive the typestate automaton into a final state. A special case of a safe region would be a region that does not change the automaton's state at all—an "identity region". For regions that are safe but no identity regions, the authors summarize the effect of this region and change the program under test to update the typestate with the region's effects all at once when the region is entered. This has the advantage that the analyzed program will execute faster because it will execute fewer transitions at runtime. However, unlike our approach, the author's analysis does not aid programmers who wish to inspect their code manually. The fact that the author's transformation changes the points at which transitions occur makes it even harder for programmers to manually inspect these program points. Dwyer and Purandare's approach is, although hybrid, not based on shadow histories and hence we have no reason to believe that it is unsound. The approach cannot generally handle groups of multiple interacting objects.

Summary. To summarize, to the best of our knowledge, our work is the first to present the notion continuation-equivalent states and an efficient algorithm to computing such equivalencies. Further, our algorithm is the first sound static-analysis algorithm that supports combinations of multiple objects and residual runtime monitors.

7. CONCLUSION

We have presented a precise flow-sensitive and partially context-sensitive typestate analysis that can handle typestate specifications that refer to multiple interacting objects, and which generates residual runtime monitors. Using an additional backwards pass, the analysis computes classes of equivalent states and disables transitions between such states. This two-pass approach allows for precise local reasoning directly at relevant program points. Although the analysis is lightweight and efficient, it is also precise, exactly telling apart property-violating program locations from irrelevant locations in more than half of the cases.

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