Generation of All Counter-Examples for Push-Down Systems

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Abstract. We present a new, on-the-fly algorithm that, given a push-down system model representing a sequential program with (recursive) procedure calls and an extended finite-state automaton representing (the negation of) a safety property, produces a succinct, symbolic representation of *all* counter-examples, i.e., traces of system behaviors that violate the property. The class of what we call *minimum-recursion loop-free counter-examples* can then be generated from this representation on an as-needed basis and presented to the user. Our algorithm is also applicable, without modification, to finite-state system models. Simultaneous consideration of multiple counter-examples can minimize the number of model checking runs needed to recognize common root causes of property violations. We illustrate the use of our techniques via application to a Java-Tar utility and an FTP-server program, and discuss a prototype tool implementation which offers several abstraction techniques for easy-viewing of generated counter-examples.

1 Introduction

Model checking [18, 5] is a widely used technique for verifying whether a system specification possesses a correctness property expressed in temporal logic. If the specification fails to satisfy the property, a counter-example in the form of a sequence of events leading to the violation of the property is produced. The sequence can then be analyzed to determine if it is a valid counter-example, or is due to an imprecise/erroneous specification of the system or property. Such an event sequence is sometimes referred to as an *error trace*.

Counter-examples play a prominent role in the recently developed technique of *abstraction refinement* [1, 11, 12, 7, 17, 16, 19, 6]. In this setting, abstract models of system specifications are used in the model-checking process, as the concrete specifications tend to be intractably large or even infinite. The counter-examples generated, however, may be infeasible in the concrete model, and hence the need for refinement of the abstract model.

In abstraction refinement, the abstract model is typically refined in an iterative fashion by considering one counter-example at a time. We argue, however, that such an approach may be overly simplified and the refinement process can be accelerated by accounting for *all* counter-examples at once. In particular, it is often the case that a

```
main() {
...
1. seteuid(root);
2. if (...)
3.     read;
   else
4.     write;
...
  }
```

Policies

1. ¬(read ∨ write)
2. read ∧ write *with* preceding seteuid(root)
3. read *with* preceding seteuid(root) ∧ ¬write
4. write *with* preceding seteuid(root) ∧ ¬read

Fig. 1. (a) Example program. (b) Policies to verify.

shared sub-sequence of multiple counter-examples reveals a common root cause for refinement.

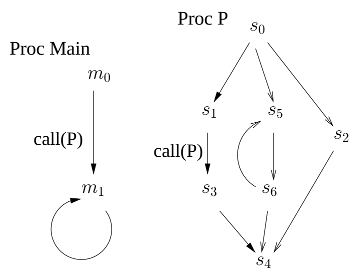
To elucidate this point, consider the problem of detecting security-policy violations in system models [21]. In this context, the generation of all counter-examples can expedite the determination of the extent of the security violations and the subsequent refinement of the policies where appropriate. Note that here the policy is refined rather than the model as in traditional abstraction refinement. For example, consider the code fragment of Figure 1(a) and the policy 1 of Figure 1(b), the latter of which disallows any read or write operations. Model checking the code against the policy reveals that the policy is violated and two traces are generated: [1, 2, 3] and [1, 2, 4]. Examination of these counter-examples reveals a common pattern—program point 1 is visited in both traces—and by refining the policy to allow reading and writing only when preceded by a `seteuid(root)` system call (policy 2 in Figure 1(b)), the desired policy is attained and no further counter-examples are produced.

Note that it was the simultaneous consideration of all counter-examples that made it possible to carry out the requisite policy refinement in one step. If counter-examples were considered one at a time, multiple steps (two in this case) would have been needed to reach the desired refinement. The intermediate policies in such an iterative process, policies 3 or 4, are given in Figure 1(b).

Most of the proposed approaches to abstraction refinement are targeted to finite-state system models. A natural question that arises then is: How are counter-examples generated in the case of *push-down systems* (PDSs) and what form should these counter-examples take? Since push-down models are a natural representation of sequential programs and play a crucial role in the treatment of potentially infinite state spaces that arise in validating recursive programs, the importance of this question becomes apparent. Although model checking of PDSs is an active area of research [4, 10, 8, 9], the topic of counter-example generation for push-down system models has gone largely uninvestigated till now.

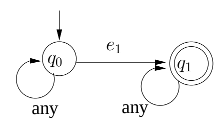
In this paper, we address the issues highlighted above by virtue of a new algorithm for the automatic generation of *all* counter-examples in a push-down system. Our main contributions can be summarized as follows.

1. We introduce the notion of *minimum-recursion loop-free* (MRLF) counter-examples for the reachability analysis of push-down systems (Section 3). An MRLF counter-example constitutes a finite representation of a potentially infinite sequence of state transitions in a PDS and assumes the form of a sequence of control-location/stack-contents pairs. The minimum-recursion loop-free aspect assures that these counter-examples do not reflect any "unnecessary" recursive procedure calls, and are thus as "succinct" as possible in a very precise sense of the word. To the best of our knowledge, this is the first proposal for PDS counter-examples to appear in the literature.
2. We present a new two-phase algorithm that, given a PDS and an extended finite-state automaton (EFSA) representing the negation of a safety property, automatically generates *all* relevant counter-examples of the property (Section 4). The first phase of the algorithm operates in an on-the-fly, property-driven fashion to generate a succinct, directed-graph representation of all possible error paths in the PDS-EFSA product, while the second phase deploys a simple stack-content-guided backward reachability analysis to construct actual MRLF counter-examples as needed.
3. The algorithm utilizes a mixture of symbolic and concrete representations. In particular, the graph representation produced by phase 1 uses regular expressions to symbolically encode potentially infinite *sets* of PDS-EFSA state-pairs. Moreover, the calculation of MRLF paths in phase 2 of the algorithm uses a symbolic encoding of the PDS transition relation to determine the set of states the EFSA can be in as a result of a recursive procedure call. Phase 2 generates MRLF counter examples by projecting abstract traces in the symbolic representation of the PDS-EFSA product onto the concrete PDS model.
4. The algorithm also utilizes a notion of a *frontier* of final states in the PDS-EFSA product: the execution paths leading to the set of final states reachable from an initial state without visiting any other final states. Constraining the generated MRLF counter-examples to fall within this frontier minimizes the number of violating paths presented to the user while still providing enough information for the comprehension of common patterns among counter-examples.
5. The algorithm has been carefully designed and implemented so that it is applicable to finite-state system models as well as PDS models (Section 5). Finite-state system models are treated by the algorithm as degenerate PDS models in which the stack depth is always equal to one.
6. Our algorithm is applicable to properties augmented with state variables. State variables are needed, for example, to identify security-critical system behaviors, specifically those related to system- and/or method-call arguments. Extended finite-state automata are used to represent properties augmented with state variables.
7. We have developed a prototype implementation of our counter-example-generation algorithm and a sophisticated graphical user interface that utilizes several abstraction techniques to render the counter-examples more comprehensible to the user (Section 6). These include selective counter-example generation, folding and unfolding of counter-examples based on a notion of "interesting event", and selective viewing of counter-example traces.
8. We have applied our techniques to a number of real-life systems including a Java-Tar utility and an FTP-server program (Section 7). A push-down model was used



$$
\begin{aligned}
P &= \{.\} \\
\Gamma &= \{m_0, m_1, s_0, s_1, \\
&\quad s_2, s_3, s_4, s_5, s_6\} \\
\Delta &= \{\langle m_0\rangle \hookrightarrow \langle s_0 m_1\rangle, \\
&\quad \langle m_1\rangle \hookrightarrow \langle m_1\rangle, \\
&\quad \langle s_0\rangle \hookrightarrow \langle s_1\rangle, \\
&\quad \langle s_1\rangle \hookrightarrow \langle s_0 s_3\rangle, \\
&\quad \langle s_3\rangle \hookrightarrow \langle s_4\rangle, \\
&\quad \langle s_4\rangle \hookrightarrow \langle \epsilon\rangle, \\
&\quad \langle s_0\rangle \hookrightarrow \langle s_5\rangle, \\
&\quad \langle s_5\rangle \hookrightarrow \langle s_6\rangle, \\
&\quad \langle s_6\rangle \hookrightarrow \langle s_5\rangle, \\
&\quad \langle s_6\rangle \hookrightarrow \langle s_4\rangle, \\
&\quad \langle s_0\rangle \hookrightarrow \langle s_2\rangle, \\
&\quad \langle s_2\rangle \hookrightarrow \langle s_4\rangle\}
\end{aligned}
$$

(a)



(b)

Fig. 2. (a) Control-flow graph and the corresponding PDS. (b) EFSA.

in the former application and a finite-state model in the latter, thereby illustrating our techniques on both kinds of system models.

2 Preliminaries

Push-down system (PDS). A PDS is a triple $\mathcal{P} = (P, \Gamma, \Delta)$ where P is a finite set of *control locations*, Γ is a finite set of *stack alphabets* and $\Delta \subseteq (P \times \Gamma) \times (P \times \Gamma^*)$ is a finite set of *transition rules*. We shall use γ, γ', ... to denote elements of Γ and u, v, w, ... to denote elements of Γ^*. We write $\langle p, \gamma\rangle \hookrightarrow \langle p', w\rangle$ to mean that $((p, \gamma), (p', w)) \in \Delta$. We restrict ourselves to PDSs where for every rule $\langle p, \gamma\rangle \hookrightarrow \langle p', w\rangle,\ |w| \leq 2$; any PDS can be put into this form with at most a linear increase in size.

A *state* of $\mathcal{P}$ is a pair $\langle p, w\rangle$ where $p \in P$ is a control location and $w \in \Gamma^*$ is a stack content. If $\langle p, \gamma\rangle \hookrightarrow \langle p', w\rangle$, then $\forall v \in \Gamma^*$ the state $\langle p', wv\rangle$ is an immediate successor of $\langle p, \gamma v\rangle$. We say that $\langle p, \gamma v\rangle$ has a transition to $\langle p', wv\rangle$ and denote it by $\langle p, \gamma v\rangle \rightarrow \langle p', wv\rangle$. A *run* of $\mathcal{P}$ is a sequence of the form $\langle p_0, w_0\rangle, \langle p_1, w_1\rangle, \ldots, \langle p_n, w_n\rangle, \ldots$ where $\langle p_i, w_i\rangle \rightarrow \langle p_{i+1}, w_{i+1}\rangle$ for all $i \geq 0$. A run can be finite or infinite.

In modeling the execution of a program, the stack content can be used to encode execution snapshots, whereas the control location captures valuations of global variables. The top of the stack points to the current program position, and the rest of the stack records the return positions of all unfinished procedure calls. Each transition rule thus specifies a change at the top of the stack. The set of control locations degenerates into a singleton set $\{.\}$ when procedures do not share variables. Henceforth, we shall consider PDSs with P a singleton set. Figure 2(a) depicts one such PDS; since there is only one control location, it is omitted from the transition rules.

Extended Finite State Automata (EFSA). An EFSA is an 8-tuple $\mathcal{E} = (Q, Q_0, F, X, E, G, A, T)$, where Q is a finite set of *states* augmented with *variables* in X, $Q_0 \subseteq$

Q is the set of *start states*, $F \subseteq Q$ is the set of *final states*, E is the set of *events*, G is a set of *boolean conditions* defining equality or dis-equality constraints over elements in X, A is a set of *assignments* which are of the form $x := y$ with $x, y \in X$, and $T \subseteq Q \times (E \times G \times A) \times Q$ is a set of *transitions*. A transition $(q, (e, g, a), q')$ is *enabled* if in state q the event $e \in E$ is present and the boolean condition $g \in G$ is satisfied. The EFSA then executes $a \in A$ and moves from state q to the next state q'. We write $q \xrightarrow{e,g,a} q'$ to denote that $(q, (e, g, a), q') \in T$. Finally, a sequence of events is *accepted* by an EFSA if it corresponds to a sequence of transitions from a start state to a final state.

To illustrate the definition, Figure 2(b) depicts the EFSA corresponding to the 8-tuple $(\{q_0, q_1\}, \{q_0\}, \{q_1\}, \{\}, \{e_1, any\}, \{tt\}, \{\}, \{q_0 \xrightarrow{any,tt,.} q_0, q_0 \xrightarrow{e_1,tt,.} q_1, q_1 \xrightarrow{any,tt,.} q_1\})$, where the event *any* is the "wildcard event" (matches any event). This EFSA accepts the sequence consisting of the event e_1.

Reachability Analysis. Given a PDS $\mathcal{P}$ and a reachability property expressed as EFSA $\mathcal{E}$, let $\lambda \subseteq P \times \Gamma \times E$ be a relation that identifies an event in E of EFSA $\mathcal{E}$ based on the control location and the top stack element of the PDS $\mathcal{P}$. Reachability analysis is performed by computing the product of $\mathcal{P}$ and $\mathcal{E}$ and then checking for an accepting path in the product. We require that, for all $p \in P$ and $\gamma \in \Gamma$, $(p, \gamma, any) \in \lambda$. The product is a push-down automaton (PDA) $\mathcal{PE}$ defined as follows:

$\mathcal{PE} = (P_{\mathcal{PE}}, P_0, \Gamma_{\mathcal{PE}}, \Delta_{\mathcal{PE}}, G_{\mathcal{PE}})$ where

1. $P_{\mathcal{PE}} \subseteq (P \times Q)$ is the set of *control locations*
2. $P_0 \subseteq P_{\mathcal{PE}}$ is the set of *initial control locations* $(P_0 \subseteq (P \times Q_0))$
3. $\Gamma_{\mathcal{PE}} = \Gamma$
4. $\Delta_{\mathcal{PE}} = \{\langle (p, q), \gamma \rangle, \langle (p', q'), w \rangle \mid \langle p, \gamma \rangle \hookrightarrow \langle p', w \rangle, q \xrightarrow{e,g,a} q', (p, \gamma, e) \in \lambda$, and g evaluates to true in $q\}$
5. $G_{\mathcal{PE}} = \{(p, q) \mid q \in F\}$

It follows that a state of $\mathcal{PE}$ is a pair $\langle p, w \rangle$, where $p \in P_{\mathcal{PE}}$ is a control location and $w \in \Gamma^*_{\mathcal{PE}}$ is a stack content. Let, $I_{\mathcal{PE}} \subseteq \{\langle p, w \rangle \mid p \in P_0, w \in \Gamma^*_{\mathcal{PE}}\}$ be the set of *intial* states and $F_{\mathcal{PE}} = \{\langle p, w \rangle \mid p \in G_{\mathcal{PE}}, w \in \Gamma^*_{\mathcal{PE}}\}$ be the set of *final* states of $\mathcal{PE}$. As previously mentioned, without loss of generality we consider PDSs with a singleton set P. Henceforth, by control locations in the product $\mathcal{PE}$, we mean states in EFSA $\mathcal{E}$. An *accepting path* in $\mathcal{PE}$ is a run along which a state in $F_{\mathcal{PE}}$ is visited.

To illustrate the definition, Figure 3(a) presents the PDA corresponding to the product of the PDS of Figure 2(a) and the EFSA of Figure 2(b), where $(., s_3, e_1) \in \lambda$. The only initial state is $\langle q_0, m_0 \rangle$. Any state with q_1 as its control location is a final state.

3 Minimum-Recursion Loop-Free Counter Examples

Given a PDS $\mathcal{P}$ and an EFSA $\mathcal{E}$ representing the negation of an invariant property φ, the existence of an accepting path in the product $\mathcal{PE}$ implies that $\mathcal{P}$ does not satisfy the invariant. A run in $\mathcal{PE}$ from an initial state to a final state is referred to as a *witness trace*. Counter-examples are obtained by projecting witness traces to the PDS as follows. Given a witness trace $\langle q_1, w_1 \rangle, \langle q_2, w_2 \rangle, \ldots \langle q_n, w_n \rangle$, the corresponding counter-example is $\langle w_1 \rangle, \langle w_2 \rangle, \ldots \langle w_n \rangle$.

(a)

$$\begin{aligned}
P_{\mathcal{PE}} = \ & \{q_0, q_1\}, P_0 = \{q_0\} \\
\Gamma_{\mathcal{PE}} = \ & \{m_0, m_1, s_0, s_1, s_2, s_3, s_4, s_5, s_6\}, G_{\mathcal{PE}} = \{q_1\} \\
\Delta_{\mathcal{PE}} = \ & \{\langle q_0, m_0\rangle \hookrightarrow \langle q_0, s_0 m_1\rangle, \langle q_1, m_0\rangle \hookrightarrow \langle q_1, s_0 m_1\rangle, \langle q_0, m_1\rangle \hookrightarrow \langle q_0, m_1\rangle, \\
& \langle q_1, m_1\rangle \hookrightarrow \langle q_1, m_1\rangle, \langle q_0, s_0\rangle \hookrightarrow \langle q_0, s_1\rangle, \langle q_0, s_0\rangle \hookrightarrow \langle q_0, s_2\rangle, \\
& \langle q_1, s_0\rangle \hookrightarrow \langle q_1, s_1\rangle, \langle q_1, s_0\rangle \hookrightarrow \langle q_1, s_2\rangle, \langle q_0, s_1\rangle \hookrightarrow \langle q_0, s_0 s_3\rangle, \\
& \langle q_1, s_1\rangle \hookrightarrow \langle q_1, s_0 s_3\rangle, \langle q_0, s_3\rangle \hookrightarrow \langle q_1, s_4\rangle, \langle q_0, s_3\rangle \hookrightarrow \langle q_0, s_4\rangle, \\
& \langle q_1, s_3\rangle \hookrightarrow \langle q_1, s_4\rangle, \langle q_0, s_2\rangle \hookrightarrow \langle q_0, s_4\rangle, \langle q_1, s_2\rangle \hookrightarrow \langle q_1, s_4\rangle, \\
& \langle q_0, s_4\rangle \hookrightarrow \langle q_0, \epsilon\rangle, \langle q_1, s_4\rangle \hookrightarrow \langle q_1, \epsilon\rangle, \langle q_0, s_0\rangle \hookrightarrow \langle q_0, s_5\rangle, \langle q_0, s_5\rangle \hookrightarrow \langle q_0, s_6\rangle, \\
& \langle q_0, s_6\rangle \hookrightarrow \langle q_0, s_5\rangle, \langle q_0, s_6\rangle \hookrightarrow \langle q_0, s_4\rangle, \langle q_1, s_0\rangle \hookrightarrow \langle q_1, s_5\rangle, \langle q_1, s_5\rangle \hookrightarrow \langle q_1, s_6\rangle, \\
& \langle q_1, s_6\rangle \hookrightarrow \langle q_1, s_5\rangle, \langle q_1, s_6\rangle \hookrightarrow \langle q_1, s_4\rangle\} \\
I_{\mathcal{PE}} = \ & \{\langle q_0, m_0\rangle\} \\
F_{\mathcal{PE}} = \ & \{\langle q_1, w\rangle \ \mid w \in \Gamma^*_{\mathcal{PE}}\}
\end{aligned}$$

(b)

$$\begin{aligned}
& (q_0, s_0, q_0), (q_0, s_0, q_1), (q_1, s_0, q_1), (q_1, s_0, q_1), (q_0, s_2, q_0), (q_1, s_2, q_1), (q_0, s_1, q_0), \\
& (q_0, s_1, q_1), (q_1, s_1, q_1), (q_0, s_3, q_0), (q_0, s_3, q_1), (q_1, s_3, q_1), (q_0, s_4, q_0), (q_1, s_4, q_1) \\
& (q_0, s_5, q_0), (q_0, s_6, q_0), (q_1, s_5, q_1), (q_1, s_6, q_1),
\end{aligned}$$

Fig. 3. (a) Product automaton. (b) Corresponding erase relation.

Note that, due to unbounded recursion, there can be an infinite number of final states in $\mathcal{PE}$ corresponding to an infinite number of stack configurations. Hence, there can also be an infinite number of witness traces. The notion of *minimum-recursion* can be used to isolate the finite subset of witness traces that do not contain any *unnecessary* recursive procedure calls.

Definition 1 (Minimum-recursion loop-free witness trace) *A witness trace in $\mathcal{PE}$, $\langle q_1, w_1\rangle$, $\langle q_2, w_2\rangle$, ..., $\langle q_n, w_n\rangle$ is* minimum-recursion loop-free *if the following conditions hold:*

1. *$\forall i$ states $\langle q_i, w_i\rangle$ appears only once*
2. *for consecutive states $\langle q_i, w_i\rangle = \langle q, \gamma w\rangle$ and $\langle q_{i+1}, w_{i+1}\rangle = \langle q', \gamma'\gamma'' w\rangle$, if $w = u\gamma'' v$ with $\gamma'' \notin u$ and $\langle q_j, w_j\rangle = \langle q, \gamma v\rangle$, $\langle q_{j+1}, w_{j+1}\rangle = \langle q', \gamma'\gamma'' v\rangle$ are consecutive states in the run with $j < i$, then for consecutive states $\langle q_k, w_k\rangle = \langle q, \gamma u_2\rangle$ and $\langle q_{k+1}, w_{k+1}\rangle = \langle q', \gamma'\gamma'' u_2\rangle$ with $k \leq j$, $w = u_1\gamma'' u_2$, $v = v_1 u_2$, $\gamma'' \notin u_2$, $\exists q_l, q_m, q_n.\ q_m \neq q_n$, such that there are runs in $\mathcal{PE}$ from $\langle q_l, \gamma'' u_1 \gamma''\rangle$ to $\langle q_m, \epsilon\rangle$ and from $\langle q_l, \gamma'' v_1\rangle$ to $\langle q_n, \epsilon\rangle$*

Condition 1 ensures that the witness trace is cycle-free. Condition 2, for minimum-recursion along a path, is more involved. First, we identify a "call-transition" where the calling location (γ) is replaced by the start location (γ') of the called procedure and the return location (γ'') in the callee. Such a call-transition is already present in the sequence if the rest of the stack content $w = u\gamma'' v$ and there are two consecutive visited states $\langle q, \gamma v\rangle$ and $\langle q', \gamma'\gamma'' v\rangle$. This represents recursive procedure calls. Note that we have identified two consecutive recursive calls of the same procedure by considering $\gamma'' \notin u$. The stack content $\gamma'' u_1 \gamma''$ denotes the increase in the stack from the first (at state $\langle q_{k+1}, w_{k+1}\rangle$) to the current (at state $\langle q_{i+1}, w_{i+1}\rangle$) recursive call of the procedure.

Similarly, $\gamma'' v_1$ denotes the increase in the stack from the first (at state $\langle q_{k+1}, w_{k+1}\rangle$) to the previous (at state $\langle q_{j+1}, w_{j+1}\rangle$) recursive call. Finally, we identify for an EFSA state q_l, $\mathcal{PE}$ runs of the form $\langle q_l, \gamma'' u_1 \gamma''\rangle$ to $\langle q_m, \epsilon\rangle$ and $\langle q_l, \gamma'' v_1\rangle$ to $\langle q_n, \epsilon\rangle$ such that $q_m \neq q_n$.

Intuitively, the set of q_m's and the set of q_n's denote the changes in EFSA state from q_l if no further recursion of the current procedure is allowed after the state $\langle q_{i+1}, w_{i+1}\rangle$ and $\langle q_{j+1}, w_{j+1}\rangle$, respectively. We refer to these changes as the *effect* of the recursion on q_l. As $\exists q_m, q_n . q_m \neq q_n$, i.e., the effect of recursions on q_l are not identical, the current recursion at $\langle q_{i+1}, w_{i+1}\rangle$ is allowed. On the other hand, if for all EFSA states, the effect of two consecutive recursions of a procedure is identical, then the last recursion is deemed unnecessary. This implies that the set of EFSA states associated with the states reachable from $\langle q_i, w_i\rangle$ is equal to the set of EFSA states associated with the states reachable from $\langle q_j, w_j\rangle$ without making any further recursions.

To illustrate the definition of minimum-recursion, consider the example of Figure 2, the corresponding product in Figure 3(a), and the witness trace $W = \langle q_0, m_0\rangle$, $\langle q_0, s_0 m_1\rangle$, $\langle q_0, s_1 m_1\rangle$, $\langle q_0, s_0 s_3 m_1\rangle$, $\langle q_0, s_1 s_3 m_1\rangle$, $\langle q_0, s_0 s_3 s_3 m_1\rangle$, $\langle q_0, s_2 s_3 s_3 m_1\rangle$, $\langle q_0, s_4 s_3 s_3 m_1\rangle$, $\langle q_0, s_3 s_3 m_1\rangle$, $\langle q_1, s_4 s_3 m_1\rangle$. W contains consecutive recursive calls to procedure P at states $\langle q_0, s_0 s_3 m_1\rangle$ and $\langle q_0, s_0 s_3 s_3 m_1\rangle$, and the effect of these calls on EFSA states q_0 and q_1 is identical. W is therefore not a minimum-recursion witness trace.

Computing the Effect. Given an EFSA state q_1 at program point γ_1 of a procedure, we would like to compute the state q_2 the EFSA will be in when the procedure exits. This is achieved by the least-model computation of the relation erase defined as follows [3]:

1. $(q_1, \gamma_1, q_2) \in \mathsf{erase}$ if $\langle q_1, \gamma_1\rangle \hookrightarrow \langle q_2, \epsilon\rangle$
2. $(q_1, \gamma_1, q_2) \in \mathsf{erase}$ if $\langle q_1, \gamma_1\rangle \hookrightarrow \langle q, \gamma\rangle$ and $(q, \gamma, q_2) \in \mathsf{erase}$
3. $(q_1, \gamma_1, q_2) \in \mathsf{erase}$ if $\langle q_1, \gamma_1\rangle \hookrightarrow \langle q, \gamma\gamma_2\rangle$, $(q, \gamma, q') \in \mathsf{erase}$ and $(q', \gamma_2, q_2) \in \mathsf{erase}$

The three rules apply respectively to the cases where $\mathcal{PE}$ makes a transition exiting, within, and entering a procedure from the state $\langle q_1, \gamma_1\rangle$. Figure 3(b) gives the erase relation corresponding to the product automaton of Figure 3(a).

Predicate Erase lifts erase from $\Gamma_{\mathcal{PE}}$ to $\Gamma^*_{\mathcal{PE}}$ as follows:

$$(q_1, \gamma w, q_2) \in \mathsf{Erase} \text{ if } (q_1, \gamma, q) \in \mathsf{erase} \wedge (q, w, q_2) \in \mathsf{Erase}$$

Tuple $(q_1, w, q_2) \in \mathsf{Erase}$ implies that there exists a run from $\langle q_1, w\rangle$ to $\langle q_2, \epsilon\rangle$, which effectively computes the effect of w on q_1. Referring back to the above example, tuples (q_0, s_3, q_0), (q_0, s_3, q_1), (q_1, s_3, q_1) and $(q_0, s_3 s_3, q_0)$, $(q_0, s_3 s_3, q_1)$, $(q_1, s_3 s_3, q_1)$ are in Erase. This implies that W is not minimum-recursion as discussed above.

Notion of Frontier Set. The *frontier* of $\mathcal{PE}$ final states is defined to be the set of witness traces leading to final states reachable from an initial state without visiting any other final states. By considering only minimum-recursion witness paths that fall within this frontier, we can further constrain the collection of counter-examples that should be presented to the user.

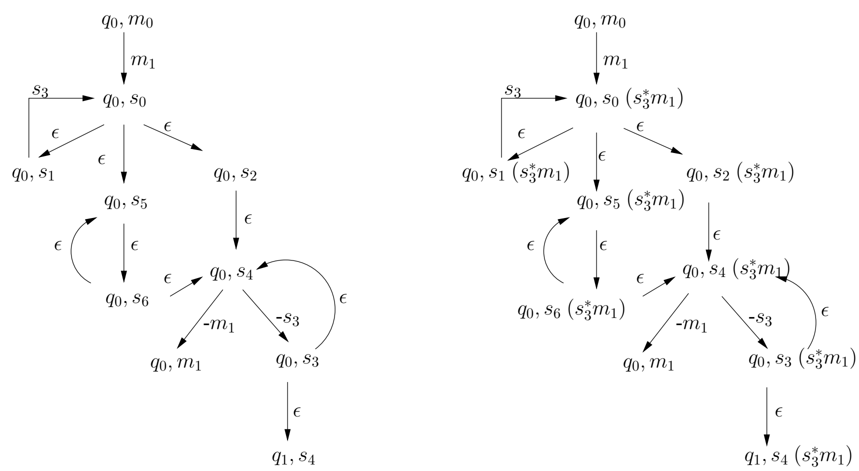
4 Algorithm for Counter-Example Generation

Our algorithm for counter-example generation proceeds in two phases. The first phase constructs an abstract graph representation, the B-graph, of all possible witness traces. Each node in the graph is a pair consisting of a control location and the top element of the stack, annotated by a regular expression defining the rest of the stack content. That is, a node in the B-graph encodes a set of states of the PDS-EFSA product that share the same control location and top element of the stack. B-graphs are constructed in a goal-directed fashion where exploration ceases upon visiting final states of the PDS-EFSA product. In phase 2, the B-graph is traversed starting from the nodes encoding the final states and minimum-recursion loop-free witness traces are constructed. We should use node and state interchangeably when there is no ambiguity, except we need to distinguish concrete states encoded by the same node.

B-graph construction proceeds in two steps. In the first step, the transition rules of $\mathcal{PE}$ are used to generate an intermediate graph called the ***rule-graph***. Each node in the rule-graph is a tuple of the form (q, γ) where $q \in P_{\mathcal{PE}}$ and $\gamma \in \Gamma_{\mathcal{PE}}$. The transition relation $\leadsto$ of the rule-graph reflects changes of control locations and stack contents when $\mathcal{PE}$ moves between states: the source and destination nodes of an edge in the rule-graph signify changes in the control location and the top element of the stack, whereas the edge label captures the change in the rest-of-stack content. An edge labeled by ϵ indicates no change to the rest of the stack. An edge labeled by a stack-alphabet symbol γ indicates that γ is pushed onto the rest of the stack in the destination state when the transition is taken. Finally, an edge labeled by "$-\gamma$" indicates that the rest of stack of the destination state can be obtained by removing a leading γ from the rest of the stack of the source state. For illustrative purposes, the rule-graph corresponding to the push-down automaton of Figure 3(a) is presented in Figure 4(a).

Step 2 of the B-graph-construction procedure annotates each node of the rule-graph with a regular expression defining the corresponding rest-of-stack content. Consider a node $\langle q, \gamma\rangle$ in the B-graph. If we view the rule graph as a finite-state automaton with final state $\langle q, \gamma\rangle$, then the regular expression we seek is the one that denotes the set of strings accepted by this automaton. Special consideration, however, must be paid to the interpretation of the concatenation operator on symbols of the form $-\gamma$. In particular, we have: $\gamma.(-\gamma) = \epsilon$, $(-\gamma).\gamma = \epsilon$, $\epsilon.(-\gamma) = -\gamma$ and $(-\gamma).\epsilon = -\gamma$. For all other cases, concatenation involving $-\gamma$ is undefined. Figure 4(b) depicts the B-graph corresponding to the rule-graph of Figure 4(a).

Figure 5 contains the pseudo-code of our algorithm. Phase 1 of the algorithm runs Procedure `constructB-graph`. It first generates the rule-graph from the product transition rules. Each node is identified by a unique index from 1 to n where n is the number of nodes. In the next step, each node in the B-graph is annotated by the corresponding regular expression. Note that the edge labels for the rule-graph and the B-graph are the same. Generation of regular expressions is performed by Function `regExp` [14, Chapter 2]: the result of the function call `regExp`(i, j, k) is a regular expression representing the set of strings defined by the sequences of transitions from the node annotated by i to the node annotated by j such that the intermediate nodes are annotated by identifiers $\leq k$.



(a) (b)

$W_1 :: \langle q_0, m_0\rangle, \langle q_0, s_0 m_1\rangle, \langle q_0, s_1 m_1\rangle, \langle q_0, s_0 s_3 m_1\rangle, \langle q_0, s_5 s_3 m_1\rangle, \langle q_0, s_6 s_3 m_1\rangle, \langle q_0, s_4 s_3 m_1\rangle, \langle q_0, s_3 m_1\rangle, \langle q_1, s_4 m_1\rangle$

$W_2 :: \langle q_0, m_0\rangle, \langle q_0, s_0 m_1\rangle, \langle q_0, s_1 m_1\rangle, \langle q_0, s_0 s_3 m_1\rangle, \langle q_0, s_2 s_3 m_1\rangle, \langle q_0, s_4 s_3 m_1\rangle, \langle q_0, s_3 m_1\rangle, \langle q_1, s_4 m_1\rangle$

(c)

Fig. 4. (a) Rule-graph. (b) B-graph. (c) Witness traces.

Phase 2 of the algorithm runs Procedure `constructPaths`, which generates all minimum-recursion loop-free witness traces via a backward-reachability analysis of the B-graph. Starting from each node $\langle q, \gamma\rangle, q \in G_{\mathcal{PE}}$ in the B-graph, let r be the regular expression annotating the node, S_r the set of strings accepted by r, and $S_r^k = \{s \mid s \in S_r, |s| = k\}$. The procedure concretizes the stack content of each $\mathcal{PE}$ state represented by the node, those with minimum stack depth first, and validates their reachability from the initial state by traversing backward through the edges. That is, it starts with the set of stack contents $\{\gamma w \mid w \in S_r^{min(k)}\}$, exhaustively validates all the corresponding states before moving on to the next set of stack contents, $\{\gamma w' \mid w' \in S_r^{min(k)+1}\}$. For example, referring to the node $\langle q_1, s_4\rangle$ in Figure 4(b), the procedure first analyzes the state $\langle q_1, s_4 m_1\rangle$, then the state $\langle q_1, s_4 s_3 m_1\rangle$, and so on.

The backward reachability proceeds as follows. Given an edge going from node $\langle q, \gamma\rangle$ to node $\langle q', \gamma'\rangle$ with a label l, a concrete state $\langle q', \gamma' w\rangle$ at node $\langle q', \gamma'\rangle$, we compute the corresponding concrete state at node $\langle q, \gamma\rangle$ as:

1. $\langle q, \gamma w\rangle$ if $l = \epsilon$
2. $\langle q, \gamma v\rangle$ if $l = \gamma''$ and $w = \gamma'' v$
3. $\langle q, \gamma\gamma' w\rangle$ if $l = -\gamma'$

A trace of concrete states is constructed by depth-first traversal starting from a concrete final state leading to the initial state following the above mentioned backward reachability process.

```
contructB-graph {
 /* create Rule-Graph R :
       the transitions relations */
  i = 1;
  make-node(q_0, γ_0); setId(q_0, γ_0, i);
  /* q_0, γ_0 is an initial state of PE */
  foreach ⟨ q, γ ⟩ ⊆ P_PE × Γ_PE {
     i++; make-node(q, γ);
     setId(q, γ, i);
  }
  n = i; /* number of nodes*/
  foreach⟨ q, γ⟩ ↪ ⟨ q', γ'⟩ && q ∉ G_PE
    add transition ⟨ q, γ⟩ ~ε~> ⟨ q', γ'⟩;
  foreach⟨ q, γ⟩ ↪ ⟨ q', γ'γ''⟩ && q ∉ G_PE
    add transition ⟨ q, γ⟩ ~γ''~> ⟨ q', γ'⟩;
  foreach⟨ q, γ⟩ ↪ ⟨ q', ε⟩ && q ∉ G_PE {
    foreach ⟨ q_1, γ_1⟩ ↪ ⟨ q_2, γ'γ''⟩
       add transition ⟨ q, γ⟩ ~−γ''~> ⟨ q', γ''⟩;
  }

 /* create B-graph : annotate nodes with
   the associated regular expressions */
 foreach (q, γ) {
   i = getId(q, γ);
   setRegExp(q, γ, regExp(1, i, n));
   /* regExp returns the regular expression
      associated with node i, i.e. (q, γ)
   */
 }
}
```

```
regExp(i, j, k) {
  if exp(i, j, k, e) ∈ store then return(e);
  else {
   if (k = 0) then {
    getNode(i, q, γ); getNode(j, q', γ');
    if(q, γ) ~e~> (q', γ') then {
     insertStore(exp(i, j, k, e)); return(e);
    }
    else { /* unreachable/*
     insertStore(exp(i, j, k, ⊥)); return(⊥);
    }
   }
   else {
    e_1 = regExp(k, k, k − 1);
    e = regExp(i, j, k − 1) ∪
         regExp(i, k, k − 1).e_1^*.regExp(i, k, k − 1);
    insertStore(exp(i, j, k, e));
    return(e);
   }
  }
}

constructPaths(B-graph) {
  len = 0;
  foreach (q, γ) ∈B-graph && q ∈ G_PE {
   getExp(q, γ, r);
   foreach s ∈ S_r && length(s) = len {
    if there is no such string s then len++;
    else {
     N = findAllPaths((q, γ s), init-state);
     /* generate all the minimum-recursion loop-free paths
        and returns number of such paths to N */
     if N! = 0 then len++; else break;
    }
   }
  }
}
```

Fig. 5. Pseudo-code for construction of B-graph and counter-example traces.

A newly constructed trace may not be minimum-recursion loop-free. Whenever we construct a trace, we check whether the minimum-recursion loop-free condition is satisfied. Procedure `constructPaths` terminates when, for a particular set of strings $S_r^n, n > min(k)$, it fails to generate any minimum-recursion loop-free traces.

Figure 4(c) lists the minimum-recursion loop-free witness traces W_1 and W_2 generated from the B-graph of Figure 4(b). Projecting these traces onto the program model yields the following counter-examples: $\langle m_0 \rangle$, $\langle s_0 m_1 \rangle$, $\langle s_1 m_1 \rangle$, $\langle s_0 s_3 m_1 \rangle$, $\langle s_5 s_3 m_1 \rangle$, $\langle s_6 s_3 m_1 \rangle$, $\langle s_4 s_3 m_1 \rangle$, $\langle s_3 m_1 \rangle$, $\langle s_4 m_1 \rangle$ and $\langle m_0 \rangle$, $\langle s_0 m_1 \rangle$, $\langle s_1 m_1 \rangle$, $\langle s_0 s_3 m_1 \rangle$, $\langle s_2 s_3 m_1 \rangle$, $\langle s_4 s_3 m_1 \rangle$, $\langle s_3 m_1 \rangle$, $\langle s_4 m_1 \rangle$, respectively.

Note that both counter-examples contain the intermediate state $\langle s_1 m_1 \rangle$, a commonality that would not be evident if counter-examples were presented one at a time. Upon examination of both counter-examples, the user may decide to introduce another event e_2, $(., s_1, e_2) \in \lambda$, and refine the property so that EFSA $\mathcal{E}$ only accepts sequences along which event e_1 is observed without any preceding e_2 event. Subsequent model checking will reveal that no such sequence is present in the PDS.

Complexity. B-graph construction amounts to the computation of regular expressions for each of the nodes. This can be performed in $O((|P_{\mathcal{PE}}| \times |\Gamma_{\mathcal{PE}}|)^3)$ time, where $|P_{\mathcal{PE}}| \times |\Gamma_{\mathcal{PE}}|$ is the number of nodes in the B-graph. The process of generating all witness traces from the B-graph requires time exponential in the number of nodes in the graph. The number of possible final states (determined by the stack content) is exponential in the number of stack alphabets ($|\Gamma_{\mathcal{PE}}|$). For each such state, we consider all possible paths ($2^{|P_{\mathcal{PE}}| \times |\Gamma_{\mathcal{PE}}|}$) from the initial state making the overall complexity exponential in $|\Gamma_{\mathcal{PE}}| + (|P_{\mathcal{PE}}| \times |\Gamma_{\mathcal{PE}}|)$.

5 Counter-example Generation for Finite-State Models

The above discussions on reachability and counter-example generation for push-down systems are directly applicable to finite-state system models. A finite-state system can be viewed as a push-down system where the stack depth is always one: the top-of-stack alphabet is always replaced by a new stack alphabet leading to no change in the stack depth. The product of a finite-state system model and an EFSA can be computed in exactly the same way as described above. Likewise, the rule-graph and B-graph construction procedures do not change.

Counter-example generation becomes simpler as we do not need to consider minimum-recursion paths: there are no recursive procedure calls in finite-state systems. Therefore the only condition imposed on witness traces is that they should be loop-free.

6 Prototype Implementation

A prototype implementation of our counter-example generation algorithm along with a sophisticated graphical user interface has been carried out in XSB Prolog [22], a tabled logic-programming environment built at SUNY, Stony Brook. XSB extends Prolog-style SLD resolution with *tabled resolution*. The principal merits of this extension are that XSB terminates more often than Prolog (e.g. for all datalog programs), avoids redundant sub-computations, and computes the well-founded model of normal logic programs. These capabilities allowed us to effectively encode the transition relations of the given PDS and EFSA as logical relations, as well compute the least relation Erase defined in Section 3 and the function `regExp`(i, j, k) of Figure 5.

For displaying counter-example traces to the user, we used *XVCG* [20], a graph-viewing utility for displaying call-graphs, flow diagrams, and the like. Counter-examples are presented with annotations on both states and transitions. Each state is identified by the entire stack content at that particular program point; the user can also select a simplified view where only the top-of-stack content is shown. Transitions are labeled by `call`/`direct`/`exit` depending on whether they represent a procedure call, direct transfer, or procedure exit, respectively.

Our implementation provides several facilities that aid the user in counter-example visualization and comprehension. One basic one is to allow the user to pre-select any number of counter-examples to be generated, particularly useful when there are large numbers of counter-example traces to the same final state. The implementation will

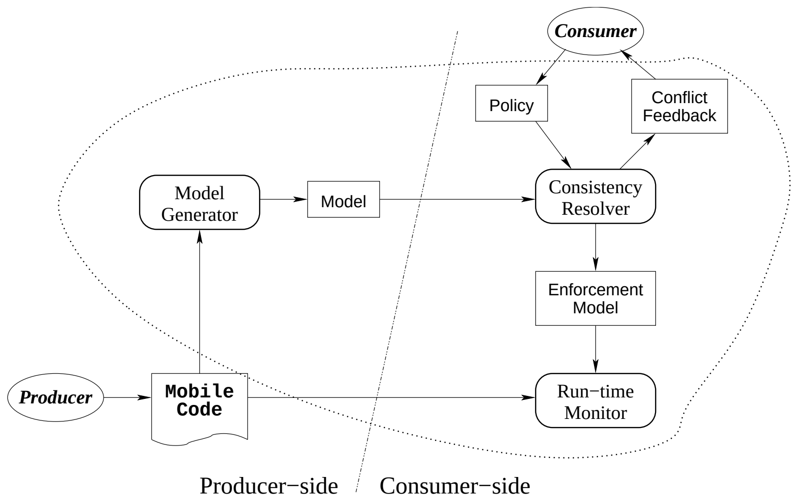


Fig. 6. Model-Carrying Code Framework.

also identify the shortest such counter-example. Further, the user is given the option of viewing any subset of generated traces.

Another feature is *counter-example abstraction*. This is performed by hiding all *uninteresting* method or system calls, thus *folding* sequences of states into a single abstract state. By uninteresting method/system calls we mean actions for which there is no state-change in the property automaton. For example, in Figure 4(b), the only interesting transition is the return from the procedure call to P; i.e., the transition from configuration $(q_0, s_3(s_3^* m_1))$ to $(q_1, s_4(s_3^* m_1))$, with concomitant state-change from q_0 to q_1 in the property automaton. Abstracted counter-examples help the user focus on the most critical aspects of violating traces. Users can also *unfold* abstracted counter-examples to view the individual steps in the trace.

Frontier sets are computed by the implementation to eliminate redundant counter-examples. Minimizing the number of counter-examples presented to the user can significantly aid overall comprehension of system behavior.

7 Case Studies

We have applied our counter-example-generation techniques to the verification of security properties in the context of the MCC (Model-Carrying Code) project [21]. Untrusted third-party code assumes many forms—including mobile code such as “active pages” (e.g. pages with Java, Javascript, VBScript, or ActiveX components), content that invokes plug-in or helper applications (e.g. document and email attachments), and software that is explicitly downloaded from a freeware or commercial site—and usually runs in user space. Assuring that it is incapable of malicious or faulty behavior is thus of paramount importance.

In the MCC approach, illustrated in Figure 6, a code producer provides an abstract model of the code's (security-relevant) behavior. Both the code and the model are sent to the consumer side, where a *consistency resolver* checks whether the model conforms to security policies selected by the consumer. When a model does not conform to a policy, the consistency resolver generates counter-examples which are presented to the consumer for further resolution, as shown in the "conflict feedback" loop in the figure. In response, the consumer can discard the code or relax the violated policy in an attempt to avoid further conflicts. Alternatively, the counter-examples may be combined with the model to produce an *enforcement model* that is given to the *runtime monitor.* The runtime monitor is responsible for confining the execution of the code so that it conforms to the enforcement model.

The consistency-resolution and conflict-feedback elements of the MCC framework naturally call for the presentation of *all* counter-examples to the user. The reasons for this are: the user will typically want to know all ways in which a security policy is violated; having knowledge of all counter-examples facilitates recognition of common root causes of counter-examples; and it allows consistency resolution to be attained in a minimal number of iterations. See also Figure 1 and its accompanying discussion in Section 1. The techniques described in this paper satisfy these requirements and were in fact designed with the MCC framework in mind.

We now describe two specific applications of our counter-example-generation techniques. The first is a publicly available tar-utility package written in Java that allows users to read and write tar archives using Java input and output streams. One important feature of the program is that if a directory is specified in the input stream, it recursively descends down the specified directory hierarchy and presents the archived files in the output stream. Such recursive behavior cannot be adequately represented by finite-state automata because of the presence of a stack. Hence, we used a push-down model to represent the behavior of the tar-utility program.

For security purposes, the code-consumer requires that the utility does not read from or write to files not specified as input arguments. The negation of this policy is specified as the EFSA given in Figure 7(a). It records in state p_2 the input arguments of the tar-utility, with `Archive` the output file and `To-Tar` the file/directory to be archived. The final state p_3 is reached via transitions labeled by actions on input or output streams with arguments that do not match `To-Tar` or `Archive`.

Model checking our push-down model of the utility reveals that the policy is violated; all counter-examples are presented in a convenient tree-like layout as illustrated in Figure 8. The counter-examples indicate that the program reads files not specified as an input argument; in particular, it requires access to dynamically loaded class files. Presentation of all counter-examples expedites recognition of this root cause of all violating paths in the system. After refining the policy to allow reading of class files, the model checker reports that the PDS model satisfies the refined policy.

The second example we considered is an ftp-server program for Linux, available from `ftp.openbsd.org`. Among other things, this program allows a remote user to execute commands such as `ls` and `cd` at the ftp server-site. Before doing so, however, the ftp-daemon must check that the user is authorized to issue these commands. This is accomplished by password-validation for which the ftp-daemon has to read from the

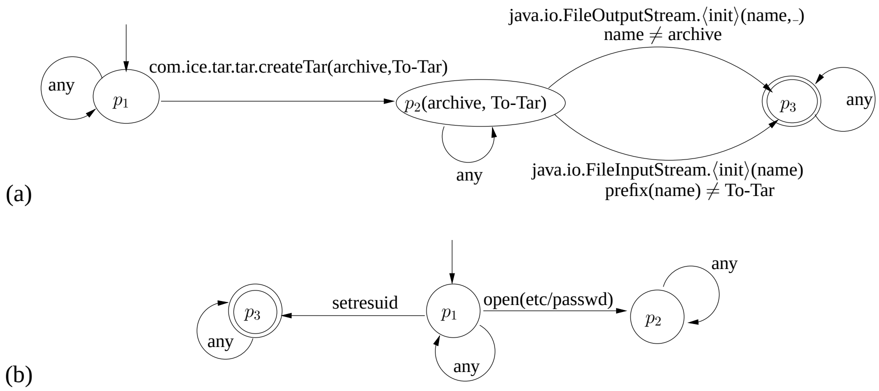


Fig. 7. (a) Tar Policy. (b) Ftp Policy.

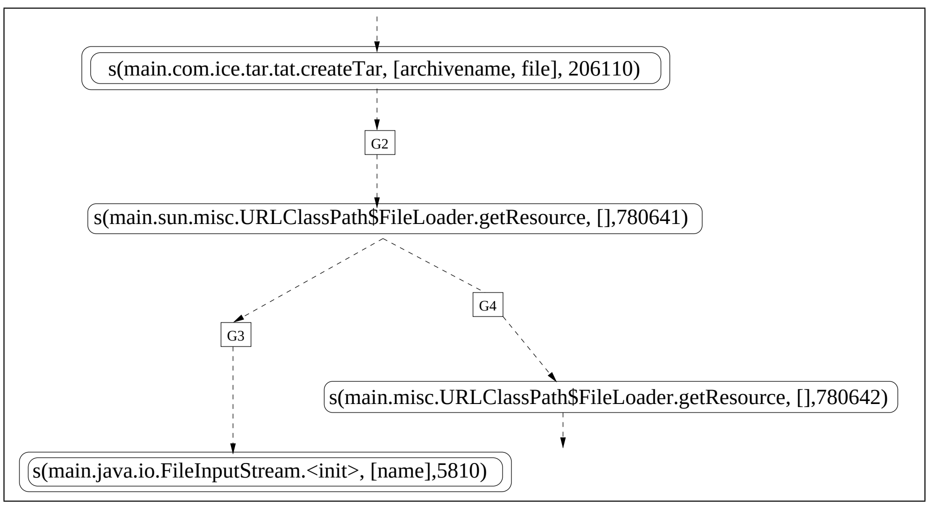


Fig. 8. Snapshot of vcg output for Tar-utility example.

`/etc/passwd` file. Once the user is successfully authenticated the ftp daemon establishes the `uid` for running such commands with `setresuid`. An interesting property to check therefore is: before issuing the system call `setresuid` there should be a call to `open(/etc/passwd)`.

We encoded the negation of this policy using the EFSA of Figure 7(b). The final state (p_3) is reachable only when there is no `open(etc/passwd)` system call preceding `setresuid`. Model checking a finite-state model of the ftp program against this policy produced numerous counter-examples, closer inspection of which revealed that the violations occurred due to abstraction performed during model generation. The abstraction in question introduced non-deterministic branch points into the model; this resulted in infeasible paths which were the source of the counter-examples.

8 Related Work

Related work in the areas of abstraction refinement and model checking of push-down systems has been discussed previously in Section 1. In other related work, [11] presents a technique where multiple counter-examples are used iteratively (one at a time) for refining an abstract system model. In [2, 13], the generation of counter-examples from the model checking of sequential-program models (control-flow graphs and labeled transition systems, respectively) is considered. The goal is to identify the *root cause* of error traces resulting from multiple executions of the model checker.

In contrast to these approaches, our algorithm produces a succinct symbolic representation of all possible witness traces of a push-down system, and subsequently generates *all* minimum-recursion loop-free counter-examples from this data structure. Collective analysis of these counter examples may then be used to identify the common root cause of security-policy violations and to refine policies where appropriate. As emphasized in [15], the presentation of all possible counter-examples minimizes the number of model-checking runs required.

9 Discussion

In this paper, we presented an algorithm that, given a push-down system model representing a sequential program with (recursive) procedure calls and an extended finite-state automaton representing (the negation of) a safety property, generates all minimum-recursion loop-free counter-examples. Our algorithm is also applicable, without modification, to finite-state system models. The utility of our techniques was illustrated via application to a Java-Tar utility and an FTP-server program, and a prototype tool implementation offering a number of abstraction techniques for easy-viewing of generated counter-examples was discussed.

As future work, we are exploring the use of machine-learning techniques to automatically identify commonality between counter-examples; the main idea is to recognize the longest common subsequence of action sequences on multiple counter-examples. Such pattern recognition enhances the usability of the tool. Another important avenue of research is to extend our techniques to the model checking of push-down systems for *liveness* properties. The major issue here is to define a finite representation for counter-examples and to extract such violating traces from model-checking results.

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