Automated Attacker Synthesis for Distributed Protocols
Expanded with Erratum

Max von Hippel, Cole Vick, Stavros Tripakis, and Cristina Nita-Rotaru
Northeastern University, Boston MA 02118, USA
{vonhippel.m, stavros, c.nitarotaru}@northeastern.edu
vick.c@husky.neu.edu

Abstract. Distributed protocols should be robust to both benign malfunction (e.g. packet loss or delay) and attacks (e.g. message replay). In this paper we take a formal approach to the automated synthesis of attackers, i.e. adversarial processes that can cause the protocol to malfunction. Specifically, given a formal threat model capturing the distributed protocol model and network topology, as well as the placement, goals, and interface (inputs and outputs) of potential attackers, we automatically synthesize an attacker. We formalize four attacker synthesis problems - across attackers that always succeed versus those that sometimes fail, and attackers that may attack forever versus those that may not - and we propose algorithmic solutions to two of them. We report on a prototype implementation called KORG and its application to TCP as a case-study. Our experiments show that KORG can automatically generate well-known attacks for TCP within seconds or minutes.

Keywords: Synthesis · Security · Distributed Protocols

1 Introduction

Distributed protocols represent the fundamental communication backbone for all services over the Internet. Ensuring the correctness and security of these protocols is critical for the services built on top of them [11]. Prior literature proposed different approaches to correctness assurance, e.g. testing [33,14], or structural reasoning [13]. Many such approaches rely on manual analysis or are ad-hoc in nature.

In this paper, we take a systematic approach to the problem of security of distributed protocols, by using formal methods and synthesis [12]. Our focus is the automated generation of attackers. But what exactly is an attack? The notion of an attack is often implicit in the formal verification of security properties: it is a counterexample violating some security specification. We build on this idea. We provide a formal definition of threat models capturing the distributed protocol model and network topology, as well as the placement, goals, and capabilities of potential attackers. Note that an attacker goal is simply the negation of a protocol property, in the sense that the goal of an attacker is to violate desirable properties that the protocol must preserve. Intuitively, an attacker is a process that, when composed with the system, results a protocol property violation.

By formally defining attackers as processes, our approach has several benefits: first, we can ensure that these processes are executable, meaning attackers are programs that reproduce attacks. This is in contrast to other approaches that generate a trace exemplifying an attack, but not a program producing the attack, e.g. [7,47]. Second, an explicit formal attacker definition allows us to distinguish different types of attackers, depending on: what exactly it means to violate a property (in some cases? in all cases?); how the attacker can behave, etc. We distinguish between ∃-attackers (that sometimes succeed in violating the security property) and ∀-attackers (that always succeed); and between attackers with recovery (that eventually revert to normal system behavior) and attackers without (that may attack forever). We make four primary contributions.

- We propose a novel formalization of threat models and attackers, where the threat models algebraically capture not only the attackers but also the attacker goals, the environmental and victim processes, and the network topology.
- We formalize four attacker synthesis problems – ∃ASP, R-∃ASP, ∀ASP, R-∀ASP – one for each of the four combinations of types of attackers.
- We propose solutions for ∃ASP and R-∃ASP via reduction to model-checking. The key idea of our approach is to replace the vulnerable processes - the victim(s) - by appropriate "gadgets", then ask a model-checker whether the resulting system violates a certain property.
– We implement our solutions in a prototype open-source tool called KORG, and apply KORG to TCP connection establishment and tear-down routines. Our experiments show KORG is able to automatically synthesize realistic, well-known attacks against TCP within seconds or minutes.

The rest of the paper is organized as follows. We present background material in Section 2. We define attacker synthesis problems in Section 3 and present solutions in Section 4. We describe the TCP case study in Section 5, present related work in Section 6, and conclude in Section 7. The solutions in Section 4 are updated to fix a mathematical error in the previously published version of this document. We explain that error in Section 8.

2 Formal Model Preliminaries

We model distributed protocols as interacting processes, in the spirit of [1]. We next define formally these processes and their composition. We also define formally the specification language that we use, namely LTL. We use $2^X$ to denote the power-set of $X$, and $\omega$-exponentiation to denote infinite repetition, e.g., $a^\omega = aaa \cdots$.

2.1 Processes

Definition 1 (Process). A process is a tuple $P = \langle AP, I, O, S, s_0, T, L \rangle$ with set of atomic propositions $AP$, set of inputs $I$, set of outputs $O$, set of states $S$, initial state $s_0 \in S$, transition relation $T \subseteq S \times (I \cup O) \times S$, and (total) labeling function $L : S \to 2^{AP}$, such that: $AP, I, O, and S$ are finite; and $I \cap O = \emptyset$.

In formal methods, Kripke Structures [27] are commonly used to describe computer programs, because they are automata (and so well-suited to describing computer programs) and their states are labeled with atomic propositions (so they are well-suited to modal logic). A process is just a Kripke Structure with inputs and outputs. Using Kripke Structures allows us to leverage LTL for free, and separating messages into inputs and outputs allows us to describe network topologies entirely using just the interfaces of the interacting processes. This idea is fundamental to our formalism of threat models in Section 4.

Let $P = \langle AP, I, O, S, s_0, T, L \rangle$ be a process. For each state $s \in S$, $L(s)$ is a subset of $AP$ containing the atomic propositions that are true at state $s$. Consider a transition $(s, x, s')$ starting at state $s$ and ending at state $s'$ with label $x$. If the label $x$ is an input, then the transition is called an input transition and denoted $s \xrightarrow{x} s'$. Otherwise, $x$ is an output, and the transition is called an output transition and denoted $s \xrightarrow{x} s'$. A transition $(s, x, s')$ is called outgoing from state $s$ and incoming to state $s'$.

A state $s \in S$ is called a deadlock iff it has no outgoing transitions. The state $s$ is called reachable if either it is the initial state or there exists a sequence of transitions $((s_i, x_i, s_{i+1}))_{i=0}^m \subseteq T$ starting at the initial state $s_0$ and ending at $s_{m+1} = s$. Otherwise, $s$ is called unreachable. The state $s$ is called input-enabled iff, for all inputs $x \in I$, there exists some state $s' \in S$ such that there exists a transition $(s, x, s') \in T$. We call $s$ an input state (or output state) if all its outgoing transitions are input transitions (or output transitions, respectively). States with both outgoing input transitions and outgoing output transitions are neither input nor output states, while states with no outgoing transitions (i.e., deadlocks) are (vacuously) both input and output states.

Various definitions of process determinism exist; ours is a variation on that of [1]. A process $P$ is deterministic iff all of the following hold: (i) its transition relation $T$ can be expressed as a (possibly partial) function $S \times (I \cup O) \to S$; (ii) every non-deadlock state in $S$ is either an input state or an output state, but not both; (iii) input states are input-enabled; and (iv) each output state has only one outgoing transition. Determinism guarantees that: each state is a deadlock, an input state, or an output state; when a process outputs, its output is uniquely determined by its state; and when a process inputs, the input and state uniquely determine where the process transitions. More intuitively, deterministic processes can be translated into concrete programs in languages like C or Java. Determinism therefore helps us make our attackers realistic.

A run of a process $P$ is an infinite sequence $r = ((s_i, x_i, s_{i+1}))_{i=0}^\infty \subseteq T^\omega$ of consecutive transitions. We use $\text{runs}(P)$ to denote all the runs of $P$. A run over states $s_0, s_1, \ldots$ induces a sequence of labels $L(s_0), L(s_1), \ldots$ called a computation. Given a (zero-indexed) sequence $\nu$, we let $\nu[i]$ denote the $i$th element of $\nu$; $\nu[i : j]$, where $i \leq j$, denote the finite suffix $\nu[i : j]$ of $\nu$, and $\nu[i : j]$ denote the infinite suffix $\nu[i : j]$ of $\nu$; we will use this notation for runs and computations.
Given two processes \( P_i = \langle AP_i, I_i, O_i, S_i, s_i^0, T_i, L_i \rangle \) for \( i = 1, 2 \), we say that \( P_1 \) is a subprocess of \( P_2 \), denoted \( P_1 \subseteq P_2 \), if \( AP_1 \subseteq AP_2, I_1 \subseteq I_2, O_1 \subseteq O_2, S_1 \subseteq S_2, T_1 \subseteq T_2 \), and, for all \( s \in S_1 \), \( L_1(s) \subseteq L_2(s) \).

### 2.2 Composition

The composition of two processes \( P_1 \) and \( P_2 \) is another process denoted \( P_1 \parallel P_2 \), capturing both the individual behaviors of \( P_1 \) and \( P_2 \) as well as their interactions with one another (e.g. Fig. 1). We define the asynchronous parallel composition operator \( \parallel \) with rendezvous communication as in [1].

**Definition 2 (Process Composition).** Let \( P_i = \langle AP_i, I_i, O_i, S_i, s_i^0, T_i, L_i \rangle \) be processes, for \( i = 1, 2 \). For the composition of \( P_1 \) and \( P_2 \) (denoted \( P_1 \parallel P_2 \)) to be well-defined, the processes must have no common outputs, i.e., \( O_1 \cap O_2 = \emptyset \), and no common atomic propositions, i.e., \( AP_1 \cap AP_2 = \emptyset \).

Then \( P_1 \parallel P_2 \) is defined below:

\[
P_1 \parallel P_2 = \langle AP_1 \cup AP_2, (I_1 \cup I_2) \setminus (O_1 \cup O_2), O_1 \cup O_2, S_1 \times S_2, (s_1^0, s_2^0), T, L \rangle
\]

... where the transition relation \( T \) is precisely the set of transitions \( (s_1, s_2) \xrightarrow{T} (s_1', s_2') \) such that, for \( i = 1, 2 \), if the label \( x \in I_i \cup O_i \) is a label of \( P_i \), then \( s_i \xrightarrow{T} s_i' \in T_i \), else \( s_i = s_i' \). \( L : S_1 \times S_2 \to 2^{AP_1 \cup AP_2} \) is the function defined as \( L(s_1, s_2) = L_1(s_1) \cup L_2(s_2) \).

Intuitively, we define process composition to capture two primary ideas: (1) rendezvous communication, meaning that a message is sent at the same time that it is received, and (2) multi-casting, meaning that a single message could be sent to multiple parties at once. We can use so-called channel processes to build asynchronous communication out of rendezvous communication (as we do in Section 5), and we can easily preclude multi-casting by manipulating process interfaces. Our definition therefore allows for a variety of communication models, making it flexible for diverse research problems. We next explain and illustrate the technical details.

A state of the composite process \( P_1 \parallel P_2 \) is a pair \( (s_1, s_2) \) consisting of a state \( s_1 \in S_1 \) of \( P_1 \) and a state \( s_2 \in S_2 \) of \( P_2 \). The initial state of \( P_1 \parallel P_2 \) is a pair \( (s_1^0, s_2^0) \) consisting of the initial state \( s_1^0 \) of \( P_1 \) and the initial state \( s_2^0 \) of \( P_2 \). The inputs of the composite process are all the inputs of \( P_1 \) that are not outputs of \( P_2 \), and all the inputs of \( P_2 \) that are not outputs of \( P_1 \). The outputs of the composite process are the outputs of the individual processes. \( P_1 \parallel P_2 \) has three kinds of transitions \( (s_1, s_2) \xrightarrow{T} (s_1', s_2') \). In the first case, \( P_1 \) may issue an output \( z \). If this output \( z \) is an input of \( P_2 \), then \( P_1 \) and \( P_2 \) move simultaneously and \( P_1 \parallel P_2 \) outputs \( z \). Otherwise, \( P_1 \) moves, outputting \( z \), but \( P_2 \) stays still (so \( s_2 = s_2' \)). The second case is symmetric to the first, except that \( P_2 \) issues the output. In the third case, \( z \) is neither an output for \( P_1 \) nor for \( P_2 \). If \( z \) is an input for both, then they synchronize. Otherwise, whichever process has \( z \) as an input moves, while the other stays still.
Fig. 1: Left is a process $P$ with atomic propositions $AP = \{p, q\}$, inputs $I = \{x, z\}$, outputs $O = \{v, w\}$, states $S = \{s_0, s_1, s_2\}$, transition relation $T = \{(s_0, w, s_0), (s_0, x, s_1), (s_0, z, s_1), (s_2, x, s_1), (s_2, v, s_2)\}$, and labeling function $L$ where $L(s_0) = \emptyset$, $L(s_1) = \{p, q\}$, and $L(s_2) = \{q\}$. Center is a process $Q$ where $L_Q(q_0) = \{r\}$ and $L_Q(q_1) = \emptyset$. Processes $P$ and $Q$ have neither common atomic propositions ($\{p, q\} \cap \{r\} = \emptyset$), nor common outputs ($\{v, w\} \cap \{x, m\} = \emptyset$), so the composition $P \parallel Q$ is well-defined. Right is the process $P \parallel Q$. Although $P \parallel Q$ is rather complicated, its only reachable states are $(s_0, q_0), (s_1, q_0)$, and $(s_1, q_1)$, and its only run is $r = \langle (s_0, q_0), (s_1, q_1), ((s_1, q_1), (s_1, q_1)) \rangle$. Non-obviously, the only computation of $P \parallel Q$ is $\sigma = \langle r \rangle, \langle p, q \rangle\rangle^\omega$.

Note that sometimes rendezvous composition is defined to match $s_1 \xrightarrow{z} s'_1$ with $s_2 \xrightarrow{z} s'_2$ to form a silent transition $(s_1, s_2) \rightarrow (s'_1, s'_2)$, but with our definition the output is preserved, so the composite transition would be $(s_1, s_2) \xrightarrow{z} (s'_1, s'_2)$. This allows for multi-casting, where an output event of one process can synchronize with multiple input events from multiple other processes. It also means there are no silent transitions.

The labeling function $L$ is total as $L_1$ and $L_2$ are total. Since we required the processes $P_1, P_2$ to have disjoint sets of atomic propositions, $L$ does not change the logic of the two processes under composition. Additionally, $\parallel$ is commutative and associative [1].

2.3 LTL

$LTL$ [35] is a linear temporal logic for reasoning about computations. In this work, we use LTL to formulate properties of processes. The syntax of LTL is defined by the following grammar:

$$\phi ::= p \mid q \mid \ldots \mid \phi_1 \land \phi_2 \mid \neg \phi_1 \mid X\phi_1 \mid U\phi_2$$

(2)

where $p, q, \ldots \in AP$ can be any atomic propositions, and $\phi_1, \phi_2$ can be any LTL formulae. Let $\sigma$ be a computation of a process $P$. If an LTL formula $\phi$ is true about $\sigma$, we write $\sigma \models \phi$. On the other hand, if $\neg(\sigma \models \phi)$, then we write $\sigma \not\models \phi$. The semantics of LTL with respect to $\sigma$ are as follows.

$$\begin{align*}
\sigma \models p & \text{ iff } p \in \sigma[0] \\
\sigma \models \phi_1 \land \phi_2 & \text{ iff } \sigma \models \phi_1 \text{ and } \sigma \models \phi_2 \\
\sigma \models \neg \phi_1 & \text{ iff } \sigma \not\models \phi_1 \\
\sigma \models X\phi_1 & \text{ iff } \sigma[1] \models \phi_1 \\
\sigma \models \phi_1 U\phi_2 & \text{ iff } (\exists K \geq 0 : \sigma[K] \models \phi_2 \text{ and } \\
& \forall 0 \leq j < K : \sigma[j] \not\models \phi_1) \\
\end{align*}$$

(3)

Essentially, $p$ holds iff it holds at the first step of the computation; the conjunction of two formulae holds if both formulae hold; the negation of a formula holds if the formula does not hold; $X\phi_1$ holds if $\phi_1$ holds in the next step of the computation; and $\phi_1 U\phi_2$ holds if $\phi_2$ holds at some future step of the computation, and until then, $\phi_1$ holds. Standard syntactic sugar include $\lor$, true, false, $F$, $G$, and $\neg$. For all LTL formulae $\phi_1, \phi_2$ and atomic
propositions $p \in \text{AP}$: $\phi_1 \lor \phi_2 \equiv \neg(\neg\phi_1 \land \neg\phi_2)$; $\text{true} \equiv p \lor \neg p$; $\text{false} \equiv \neg\text{true}$; $F\phi_1 \equiv \text{trueU}\phi_1$; $G\phi_1 \equiv \neg F\neg\phi_1$; and $\phi_1 \rightarrow \phi_2 \equiv (\neg\phi_1) \lor (\phi_1 \land \phi_2)$.

Example formulae include:

- Lunch will be ready in a moment: $X(\text{lunch-ready})$.
- I always eventually sleep: $GF\text{sleep}$.
- I am hungry until I eat: $GU\text{eat}$.
- $A$ and $B$ are never simultaneously in their crit states: $G(\neg(\text{crit}_A \land \text{crit}_B))$.

An LTL formula $\phi$ is called a safety property iff it can be violated by a finite prefix of a computation, or a liveness property iff it can only be violated by an infinite computation $[3]$. For a process $P$ and LTL formula $\phi$, we write $P \models \phi$ iff, for every computation $\sigma$ of $P$, $\sigma \models \phi$. For convenience, we naturally elevate our notation for satisfaction on computations to satisfaction on runs, that is, for a run $r$ of a process $P$ inducing a computation $\sigma$, we write $r \models \phi$ and say “$r$ satisfies $\phi$” iff $\sigma \models \phi$, or write $r \not\models \phi$ and say “$r$ violates $\phi$” iff $\sigma \not\models \phi$.

3 Attacker Synthesis Problems

We want to synthesize attackers automatically. Intuitively, an attacker is a process that, when composed with the system, violates some property. There are different types of attackers, depending on what it means to violate a property (in some cases? in all cases?), as well as on the system topology (threat model). Next, we define the threat model and attacker concepts formally, followed by the problems considered in this paper.

3.1 Threat Models

A threat model or attacker model prosaically captures the goals and capabilities of an attacker with respect to some victim and environment. Algebraically, it is difficult to capture the attacker goals and capabilities without also capturing the victim and the environment, so our abstract threat model includes all of the above. Our threat model captures: how many attacker components there are; how they communicate with each other and with the rest of the system: what messages they can intercept, transmit, etc; and the attacker goals. We formalize the concept of a threat model in what follows.

Definition 3 (Input-Output Interface). An input-output interface is a tuple $(I, O)$ such that $I \cap O = \emptyset$ and $I \cup O \neq \emptyset$. The class of an input-output interface $(I, O)$, denoted $C(I, O)$, is the set of processes with inputs $I$ and outputs $O$. Likewise, $C(P)$ denotes the input-output interface the process $P$ belongs to. (e.g. Fig. 3)

Definition 4 (Threat Model). A threat model is a tuple $(P, (Q_i)_{i=0}^m, \phi)$ where $P, Q_0, ..., Q_m$ are processes, each process $Q_i$ has no atomic propositions (i.e., its set of atomic propositions is empty), and $\phi$ is an LTL formula such that $P \parallel Q_0 \parallel ... \parallel Q_m \models \phi$. We also require that the system $P \parallel Q_0 \parallel ... \parallel Q_m$ satisfies the formula $\phi$ in a non-trivial manner, that is, that $P \parallel Q_0 \parallel ... \parallel Q_m$ has at least one infinite run.

In a threat model, the process $P$ is called the target process, and the processes $Q_i$ are called vulnerable processes. The goal of the adversary is to modify the vulnerable processes $Q_i$ so that composition with the target process violates $\phi$. (We assume that prior to the attack, the protocol behaves correctly, i.e., it satisfies $\phi$.) For example, in TM1 of Fig. 2 the target process is Alice composed with Bob, and the vulnerable processes are Oscar and Trudy, while in TM5, the target process is the composition of Jacob, Simon, Sophia, and Juan, and the only vulnerable process is Isabelle.
Mallory Bob

\[ \text{TM}_1 = (\text{Alice } \parallel \text{Bob}, (\text{Oscar, Trudy}), \phi_1) \]

Alice Trudy Bob

\[ \text{TM}_2 = (\text{Alice, (Oscar), } \phi_2) \]

Alice Oscar

\[ \text{TM}_3 = (\text{Alice } \parallel \text{Bob, (Mallory), } \phi_3) \]

Alice Mallory Bob

\[ \text{TM}_4 = (\text{Alice } \parallel \text{Mark, (Eve), } \phi_4) \]

Alice Mark

\[ \text{TM}_5 = (\text{Jacob } \parallel \text{Simon } \parallel \text{Sophia } \parallel \text{Juan, (Isabelle), } \phi_5) \]

Simon Sophia

Fig. 2: Example Threat Models. The properties \( \phi_i \) are not shown. Solid and dashed boxes are processes; we only assume the adversary can exploit the processes in the dashed boxes. \( \text{TM}_1 \) describes a distributed off-path attacker scenario, \( \text{TM}_2 \) describes a classical man-in-the-middle scenario, and \( \text{TM}_4 \) describes a one-directional man-in-the-middle, or, depending on the problem formulation, an eavesdropper. \( \text{TM}_5 \) is a threat model with a distributed victim where the attacker cannot affect or read messages from Simon to Juan. Note that a directed edge in a network topology from Node 1 to Node 2 is logically equivalent to the statement that a portion of the outputs of Node 1 are also inputs to Node 2. In cases where the same packet might be sent to multiple recipients, the sender and recipient can be encoded in a message subscript. Therefore, the entire network topology is implicit in the interfaces of the processes in the threat model according to the composition definition.

3.2 Attackers

**Definition 5 (Attacker).** Let \( \text{TM} = (P, (Q_i)_{i=0}^m, \phi) \) be a threat model. Then \( \bar{A} = (A_i)_{i=0}^m \) is called a \( \text{TM-attacker} \) if \( P \parallel A_0 \parallel ... \parallel A_m \not\models \phi \), and, for all \( 0 \leq i \leq m \): \( A_i \) is a deterministic process; \( A_i \) has no atomic propositions, and \( A_i \in C(Q_i) \).

The existence of a \( (P, (Q_i)_{i=0}^m, \phi) \)-attacker means that if an adversary can exploit all the \( Q_i \), then the adversary can attack \( P \) with respect to \( \phi \). Note that an attacker \( \bar{A} \) cannot succeed by blocking the system from having any runs at all. Indeed, \( P \parallel A_0 \parallel ... \parallel A_m \not\models \phi \) implies that \( P \parallel A_0 \parallel ... \parallel A_m \) has at least one infinite run violating \( \phi \).

Real-world computer programs implemented in languages like C or Java are called concrete, while logical models of those programs implemented as algebraic transition systems such as processes are called abstract. The motivation for synthesizing abstract attackers is ultimately to recover exploitation strategies that actually work against concrete protocols. So, we should be able to translate an abstract attacker (Fig. 5) into a concrete one (Fig. 11). Determinism guarantees that we can do this. We also require the attacker and the vulnerable processes to have no atomic propositions, so the attacker cannot “cheat” by directly changing the truth-hood of the property it aims to violate.

We can define an attacker for many properties at once by conjoining those properties (e.g. \( \phi_1 \land \phi_2 \land \phi_3 \)), or for many processes at once by composing those processes (e.g. \( P_1 \parallel P_2 \parallel P_3 \)). We therefore do not lose expressibility compared to a definition that explicitly allows many properties or processes.

For a given threat model many attackers may exist. We want to differentiate attacks that are more effective from attacks that are less effective. One straightforward comparison is to partition attackers into those that always violate \( \phi \), and those that only sometimes violate \( \phi \). We formalize this notion with \( \exists \)-attacker and \( \forall \)-attacker.

**Definition 6 (\( \exists \)-Attacker vs \( \forall \)-Attacker).** Let \( \bar{A} \) be a \( (P, (Q_i)_{i=0}^m, \phi) \)-attacker. Then \( \bar{A} \) is a \( \forall \)-attacker if \( P \parallel A_0 \parallel ... \parallel A_m \not\models \phi \). Otherwise, \( \bar{A} \) is an \( \exists \)-attacker.

A \( \forall \)-attacker \( \bar{A} \) always succeeds, because \( P \parallel \bar{A} \models \neg \phi \) means that every behavior of \( P \parallel \bar{A} \) satisfies \( \neg \phi \), that is, every behavior of \( P \parallel \bar{A} \) violates \( \phi \). Since \( P \parallel \bar{A} \not\models \phi \), there must exist a computation \( \sigma \) of \( P \parallel \bar{A} \) such that \( \sigma \models \neg \phi \), so, a \( \forall \)-attacker cannot succeed by blocking. An \( \exists \)-attacker is any attacker that is not a \( \forall \)-attacker, and every attacker succeeds in at least one computation, so an \( \exists \)-attacker sometimes succeeds, and sometimes does not. In most real-world systems, infinite attacks are impossible, implausible, or just uninteresting. To avoid such attacks, we define an attacker that produces finite-length sequences of adversarial behavior, and then “recovers”, meaning that it behaves like the vulnerable process it replaced (see Fig. 4).
Definition 7 (Attacker with Recovery). Let \( \vec{A} \) be a \((P, (Q_i)_{i=0}^m, \phi)\)-attacker. If, for each \( 0 \leq i \leq m \), the attacker component \( A_i \) consists of a finite directed acyclic graph (DAG) ending in the initial state of the vulnerable process \( Q_i \), followed by all of the vulnerable process \( Q_i \), then we say the attacker \( \vec{A} \) is an attacker with recovery. We refer to the \( Q_i \) postfix of each \( A_i \) as its recovery, as in Fig. 3, \( A_3 \). Intuitively, the vulnerable process is appended to the attacker.

Note that researchers sometimes use “recovery” to mean when a system undoes the damage caused by an attack. We use the word differently, to mean when the property \( \phi \) remains violated even under modus operandi subsequent to attack termination.

Fig. 3: From left to right: processes \( P \), \( Q \), \( A_1 \), \( A_2 \), \( A_3 \). Let \( \phi = G\{\text{OK}\} \), and let the interface of \( Q \) be \( C(Q) = (\emptyset, \{a, b, c\}) \). Then \( P \parallel Q \models \phi \). \( A_1 \) and \( A_2 \) are both deterministic and have no input states. Let \( C(A_1) = C(A_2) = C(Q) \). Then, \( A_1 \) and \( A_2 \) are both \((P, (Q), \phi)\)-attackers. \( A_1 \) is a \( \exists \)-attacker, and \( A_2 \) is an \( \forall \)-attacker. \( A_3 \) is a \( \forall \)-attacker with recovery consisting of a DAG starting at \( a_3^0 \) and ending at the initial state \( q_0 \) of \( Q \), plus all of \( Q \), namely the recovery.

Fig. 4: Suppose \( \vec{A} = (A_i)_{i=0}^m \) is attacker with recovery for \( \text{TM} = (P, (A_i)_{i=0}^m, \phi) \). Further suppose \( A_i \) has initial state \( a_i^0 \), and \( Q_i \) has initial state \( q_i^0 \). Then \( A_i \) should consist of a DAG starting at \( a_i^0 \) and ending at \( q_i^0 \), plus all of \( Q_i \), called the recovery, indicated by the shaded blob. Note that if some \( Q_i \) is non-deterministic, then there can be no attacker with recovery, because \( Q_i \) is a subprocess of \( A_i \), and all the \( A_i \)s must be deterministic in order for \( \vec{A} \) to be an attacker.

3.3 Attacker Synthesis Problems

Each type of attacker - \( \exists \) versus \( \forall \), with recovery versus without - naturally induces a synthesis problem.
Problem 1 (∃-Attacker Synthesis Problem (∃ASP)). Given a threat model TM, find a TM-attacker, if one exists; otherwise state that none exists.

Problem 2 (Recovery ∃-Attacker Synthesis Problem (R-∃ASP)). Given a threat model TM, find a TM-attacker with recovery, if one exists; otherwise state that none exists.

We defined ∃ and ∀-attackers to be disjoint, but, if the goal is to find an ∃-attacker, then surely a ∀-attacker is acceptable too; we therefore did not restrict the ∃-problems to only ∃-attackers. Next we define the two ∀-problems, which remain for future work.

Problem 3 (∀-Attacker Synthesis Problem (∀ASP)). Given a threat model TM, find a TM-∀-attacker, if one exists; otherwise state that none exists.

Problem 4 (Recovery ∀-Attacker Synthesis Problem (R-∀ASP)). Given a threat model TM, find a TM-∀-attacker with recovery, if one exists; otherwise state that none exists.

4 Solutions

We present solutions ∃ASP and R-∃ASP for any number of attackers, and for both safety and liveness properties. Our success criteria are soundness and completeness. Both solutions are polynomial in the product of the size of $P$ and the sizes of the interfaces of the $Q_i$s, and exponential in the size of the property $\phi$ [43]. For real-world performance, see Section 5.

We reduce ∃ASP and R-∃ASP to model-checking. The idea is to replace the vulnerable processes $Q_i$ with appropriate “gadgets”, then ask a model-checker whether the system violates a certain property. We prove that existence of a violation (a counterexample) implies existence of an attacker, and for the ∃ASP, that the implication goes both ways. Then, we show how to transform the counterexample into an attacker. The gadgets and the LTL formula are different, depending on whether we seek attackers without or with recovery.

4.1 Gadgetry

We begin by defining lassos and bad prefixes. A computation $\sigma$ is a lasso if it equals a finite word $\alpha$, then infinite repetition of a finite word $\beta$, i.e., $\sigma = \alpha \cdot \beta^\omega$. A prefix $\alpha$ of a computation $\sigma$ is called a bad prefix for $P$ and $\phi$ if $P$ has $\geq 1$ runs inducing computations starting with $\alpha$, and every computation starting with $\alpha$ violates $\phi$. We naturally elevate the terms lasso and bad prefix to runs and their prefixes. We assume a model checker: a procedure $MC(P, \phi)$ that takes as input a process $P$ and property $\phi$, and returns $\emptyset$ if $P \models \phi$, or one or more violating lasso runs or bad prefixes of runs for $P$ and $\phi$, otherwise [3].

Attackers cannot have atomic propositions. So, the only way for $A$ to attack TM is by sending and receiving messages, hence the space of attacks is within the space of labeled transition sequences. The Daisy Process nondeterministically exhausts the space of input and output events of a vulnerable process.

Definition 8 (Daisy Process). Let $Q = \langle \emptyset, I, O, S, s_0, T, L \rangle$ be a process with no atomic propositions. Then the daisy of $Q$, denoted DAISY$(Q)$, is the process defined below, where $L' : \{d_0\} \rightarrow \{\emptyset\}$ is the map such that $L'(d_0) = \emptyset$.

$$DAISY(Q) = \langle \emptyset, I, O, \{d_0\}, d_0, \{(d_0, w, d_0) \mid w \in I \cup O\}, L' \rangle$$ (4)
Next, we define a Daisy with Recovery. This gadget is an abstract process, i.e., a generalized process with a non-empty set of initial states $S_0 \subseteq S$. Composition and LTL semantics for abstract processes are naturally defined. We implicitly transform processes to abstract processes by wrapping the initial state in a set.

**Definition 9 (Daisy with Recovery).** Given a process $Q_i = \langle \emptyset, I, O, S, s_0, T, L \rangle$, the daisy with recovery of $Q_i$, denoted $\text{RDaisy}(Q_i)$, is the abstract process $\text{RDaisy}(Q_i) = \langle AP, I, O, S', S_0, T', L' \rangle$, with atomic propositions $AP = \{ \text{recover}_i \}$, states $S' = S \cup \{d_0\}$, initial states $S_0 = \{s_0, d_0\}$, transitions $T' = T \cup \{(d_0, x, w_0) \mid x \in I \cup O, w_0 \in S_0\}$, and labeling function $L' : S' \rightarrow 2^{AP}$ that takes $s_0$ to $\{ \text{recover}_i \}$ and other states to $\emptyset$. (We reserve the symbols $\text{recover}_0, ...$ for use in daisies with recovery, so they cannot be sub-formulae of the property in any threat model.)

The daisy with recovery gadget is illustrated in Fig. 6.

![Diagram](image-url)

Fig. 6: $\text{RDaisy}(Q_i)$ is the result of (1) directly connecting $\text{DAisy}(Q_i)$ to the initial state of $Q_i$ in every possible way, (2) allowing both the initial state of the daisy and the initial state of $Q_i$ to be initial states of the abstract process, and (3) adding a single atomic proposition $\text{recover}_i$ to the initial state of $Q_i$.

### 4.2 Solution to $\exists$ASP

Let $TM = \langle P, (Q_i)_{i=0}^{m}, \phi \rangle$ be a threat model. Our goal is to find an attacker for $TM$, if one exists, or state that none exists, otherwise. First, we check whether the system $P \parallel \text{DAisy}(Q_0) \parallel ... \parallel \text{DAisy}(Q_m)$ satisfies $\phi$. If it does, then no attacker exists, as the daisy processes encompass any possible attacker behavior. Define a set $R$ returned by the model-checker $MC$:

$$R = MC(P \parallel \text{DAisy}(Q_0) \parallel ... \parallel \text{DAisy}(Q_m), \phi)$$  \hspace{1cm} (5)$$

If $R = \emptyset$ then no attacker exists.
Solution to \( \exists A \Sigma P \) (Pseudocode)

1. \( R \leftarrow \text{MC}(P \parallel \text{DAISY}(Q_0) \parallel \ldots \parallel \text{DAISY}(Q_m), \phi) \)
2. If \( (R = \emptyset) \) then return "no TM-attacker exists"

On the other hand, if the system violates \( \phi \), then we can transform a violating run into a set of attacker processes by projecting it onto the corresponding interfaces. Choose a violating run or bad prefix \( r \in R \) arbitrarily. Either \( r = \alpha \) is some finite bad prefix, or \( r = \alpha \cdot \beta^\omega \) is a violating lasso. For each \( 0 \leq i < m \), let \( \alpha_i \) be the projection of \( \alpha \) onto the process \( \text{DAISY}(Q_i) \). That is, \( \alpha_i = [\emptyset] \); then for each \( (\bar{s}, x, \bar{s}') \in \alpha \), if \( x \) is an input or an output of \( Q_i \), and \( q, q' \) are the states \( \text{DAISY}(Q_i) \) embodies in \( \bar{s}, \bar{s}' \), add \( (q, x, q') \) to \( \alpha_i \).

3. Choose \( r \in R \) arbitrarily.
4. Either \( r = \alpha \cdot \beta^\omega \) is a lasso, or \( r = \alpha \) is a bad prefix.
5. \( (\alpha_i)^m_{i=0} \leftarrow ([\emptyset])^m_{i=0} \)
6. \( \vec{A} \leftarrow (\perp)^m_{i=0} \)
7. For \( 0 \leq i < m \) do
   7.1. For \( 0 \leq j < |\alpha_i| \) do
      7.1.1. \( (s, x, \bar{s}') \leftarrow \alpha_i[j] \)
   7.1.2. \( \text{If} \; (x \in I_i \cup O_i) \; \text{then} \; \alpha_i, \text{append}((s[i], x, s'[i])) \)

For each \( \alpha_i \), create an incomplete process \( A_i^\omega \) with a new state \( s^\alpha_{j+1} \) and transition \( s^\alpha_j \xrightarrow{a} s^\alpha_{j+1} \) for each \( \alpha_i[j] = (d_i^0, z, d_i^0) \) for \( 0 \leq j < |\alpha_i| \). If \( r = \alpha \cdot \beta^\omega \) is a lasso, then for each \( 0 \leq i < m \), define \( A_i^\beta \) from \( \beta_i \) in the same way that we defined \( A_i^\alpha \) from \( \alpha_i \); let \( A_i' \) be the result of merging the first and last states of \( A_i^\beta \) with the last state of \( A_i^\alpha \).

7.2. Create a new state \( s^\alpha_0 \)
7.3. \( S_i \leftarrow \{s^\alpha_0\} \)
7.4. \( T_i \leftarrow \emptyset \)
7.5. For \( 0 \leq j < |\alpha_i| \) do
   7.5.1. \( (z, \bar{s}) \leftarrow \alpha_i[j] \)
   7.5.2. Create a new state \( s^\alpha_{j+1} \)
   7.5.3. \( S_i, \text{append}(s^\alpha_{j+1}) \)
   7.5.4. \( T_i, \text{append}((s^\alpha_j, z, s^\alpha_{j+1})) \)
7.6. Let \( L_i : S_i \rightarrow 2^\emptyset \) denote the function \( \lambda s.\emptyset \)
7.7. \( A_i^\omega \leftarrow (\emptyset, I_i, O_i, S_i, T_i, L_i) \)
7.8. If \( (r = \alpha \cdot \beta^\omega \) is a lasso \) then
   7.8.1. Define \( \beta_i \) from \( \beta \) in the same way we defined \( \alpha_i \) from \( \alpha \) in 7.1
   7.8.2. Define \( A_i^\beta = (\emptyset, I_i, O_i, S^\beta_i, T^\beta_i, L^\beta_i) \) from \( \beta_i \) as we defined \( A_i^\alpha \) from \( \alpha_i \) in 7.2–7.6
   7.8.3. \( A_i \leftarrow A_i^\omega \text{ glued to } A_i^\beta \) where we merge \( A_i^\alpha \)’s final state \( s^\alpha_{|\alpha_i|} \) with \( A_i^\beta \)’s initial state \( s^\beta_0 \)

Otherwise, if \( r = \alpha \) is a bad prefix, let \( A_i' \) be the result of adding an input self-loop to the last state of \( A_i^\alpha \), or an output self-loop if \( Q_i \) has no inputs. Either way, \( A_i' \) is an incomplete attacker. Finally let \( A_i \) be the result of making every input state in \( A_i' \) input-enabled via self-loops, and return the attacker \( \vec{A} = (A_i)^m_{i=0} \).

7.9. Else if \( (I_i \neq \emptyset) \) then
   7.9.1. \( A_i \leftarrow \) the result of adding a single arbitrary input self-loop to the final state \( s^\alpha_{|\alpha_i|} \) of \( A_i^\alpha \)
7.10. Else \( A_i \leftarrow \) the result of adding a single arbitrary output self-loop to the final state \( s^\alpha_{|\alpha_i|} \) of \( A_i^\alpha \)
7.11. For \( s \in \text{states}(A_i) \) do
    7.11.1. \( \text{If } s \text{ is an input state then } \text{For } x \in I_i \text{ do transitions} \begin{array}{l} A_i \text{.append}((s, x, s)) \end{array} \)
8. return \( \vec{A} \)
We inductively demonstrate that if we prove 2) then 1). Suppose we need to show the result of the run into attacker. Then the following hold. 1) \( P \parallel Q_0 \parallel Q_1 = \phi \).

**Threat Model:** \( TM' = (P, (Q_0, Q_1), \phi) \), where the processes from left to right are \( P, Q_0, \) and \( Q_1 \), and where \( \phi = FG I \). Let \( P \) have inputs \( k \) and \( m \), and output \( n \). \( Q_0 \) has no inputs, and output \( m \). \( Q_1 \) has inputs \( n \) and \( h \), and output \( k \). Recall that \( P \parallel Q_0 \parallel Q_1 = \phi \).

**Sketch of Proof.** A run \( r \in R \) where \( R \) is defined as in Equation[5] returns a TM-attacker. Showing the result is \( \exists \) ASP. Example threat model TM-attacker exists. 2) If \( R \neq \emptyset \), then the procedure above eventually returns a TM-attacker.

**Application of solution:** \( r \) is projected and translated into an attacker \( A = (A_0, A_1) \).

**Violating run:** A run \( r \in R \) where \( R \) is defined as in Equation[5]

\[
r = \begin{pmatrix} p_0 \to k \parallel p_1 \to m \parallel p_2 \to m \parallel p_3: \{ l \} \parallel m \to m' \parallel q_0 \to q_0' \parallel q_1 \to q_1' \parallel \end{pmatrix}
\]

**Fig. 7:** Illustration of solution to \( \exists \) ASP. Example threat model TM' on top, followed by a violating run in \( R \), followed by translation of the run into attacker.

**Theorem 1 (\( \exists \) ASP Solution is Sound and Complete).** Let \( TM = (P, (Q_i)_{i=0}^n, \phi) \) be a threat model, and define \( R \) as in Eqn. [5] Then the following hold. 1) \( R \neq \emptyset \) iff a TM-attacker exists. 2) If \( R \neq \emptyset \), then the procedure above eventually returns a TM-attacker.

**Sketch of Proof.** We prove 2) then 1). Suppose \( r \in R \). Processes are finite and threat models are finitely large, so the procedure eventually terminates. We need to show the result \( A = (A_i)_{i=0}^n \) is a TM-attacker. Showing the result is deterministic is straightforward, and it should be equally clear that each \( A_i \) has the same interface as its respective \( Q_i \). We inductively demonstrate that \( P \parallel A_0 \parallel \ldots \parallel A_m \) has some run \( r' \) that is I/O-equivalent to the run \( r \) and induces the same computation. So then \( r' \neq \phi \), so \( A \) is a TM-attacker and therefore 2) holds.

We now turn our attention to 1). If a TM-attacker \( A' \) exists, then \( P \parallel A'_0 \parallel \ldots \parallel A'_m \) has a run \( r' \) violating \( \phi \). The daisies can do everything the \( Q_i \)s can do and more, so the daisies yield some I/O-equivalent run \( r' \) violating \( \phi \), and so \( R \neq \emptyset \). On the other hand, if \( R = \emptyset \) then we can easily prove by way of contradiction that no attacker exists, since
attackers, daisies, and vulnerable processes have no atomic propositions, and therefore any violating run of an attacker with $P$ could be translated into an I/O-equivalent run of the daisies with $P$ inducing the same computation. So 1) holds and we are done.

4.3 Solution to R-$\exists$ASP

Let $\text{TM} = (P, (Q_i)_{i=0}^m, \phi)$ be a threat model as before. Now our goal is to find a TM-attacker with recovery, if one exists, or state that none exists, otherwise. The idea to solve this problem is similar to the idea for finding attackers without recovery, with two differences. First, the daisy processes are now more complicated, and include recovery to the original $Q_i$ processes. Second, the formula used in model-checking is not $\phi$, but a more complex formula $\psi$ to ensure that the attacker eventually recovers, i.e., all the attacker components eventually recover. We define the property $\psi$ so that in prose it says “if all daisies eventually recover, then $\phi$ holds”. We then define $R$ like before, except we replace daisies with daisies with recovery, and $\phi$ with $\psi$, as defined below.

$$\psi = \left( \bigwedge_{0 \leq i \leq m} F_{\text{recover}_i} \right) \implies \phi$$

$$R = \text{MC}(P \parallel \text{RDAisy}(Q_0) \parallel \ldots \parallel \text{RDAisy}(Q_m), \psi)$$

If $R = \emptyset$ then no attacker with recovery exists. If any $Q_i$ is not deterministic, then likewise no attacker with recovery exists, because our attacker definition requires the attacker to be deterministic, but if $Q_i$ is not and $Q_i \subseteq A_i$ then neither is $A_i$.

Solution to R-$\exists$ASP (Pseudocode)

1. $R \leftarrow \text{MC}(P \parallel \text{RDAisy}(Q_0) \parallel \ldots \parallel \text{RDAisy}(Q_m), (\bigwedge_{0 \leq i \leq m} F_{\text{recover}_i} \implies \phi))$
2. If ($R = \emptyset$ or any $Q_i$ is not deterministic) then return “no TM-attacker exists”

Otherwise, choose a violating run (or bad prefix) $r \in R$ arbitrarily. We proceed as we did for $\exists$ASP but with three key differences. First, we use $\text{RDAisy}(Q_i)$ as opposed to $\text{DAisy}(Q_i)$. Second, for each $0 \leq i \leq m$, instead of using $A_i^{a}$ and $A_i^{b}$, we compute a program $A_i^{rec}$ which replays the actions taken by $\text{RDAisy}(Q_i)$ until $\text{recover}_i$ is satisfied. We glue this program to the recovery $Q_i$. Third, instead of using self-loops to input-enable input states, we use input transitions to the initial state of $Q_i$. This ensures the pre-recovery portion is a DAG. Then we return $\tilde{A} = (A_i)_{i=0}^m$.
Next we prove that our solution is sound, and that our solution is complete for the subset of attackers with recovery which eventually recover, which we refer to as terminating.

**Definition 10 (Terminating Attacker).** Let \( \vec{A} \) be a TM-attacker with recovery. If \( P \parallel A_0 \parallel \cdots \parallel A_m \) has at least one run violating \( \phi \) in which each attacker component \( A_i \) reaches the initial state \( q^0_i \) of its corresponding recovery \( Q_i \), then we say that \( \vec{A} \) is a Terminating Attacker.

In order to prove these results, we need to first prove some smaller Lemmas. Soundness means that given a threat model \( \text{TM} = (P, (Q_i)_{i=0}^m, \phi) \), if the solution returns a result \( \vec{A} \), then \( \vec{A} \) is a TM-attacker with recovery. Given some such TM and \( \vec{A} \), we show in Lemma 1 that \( \vec{A} \) is the correct size (consisting of \( m+1 \) components); in Lemma 2 that \( \vec{A} \) satisfies the syntactic requirements for an attacker; and in Lemma 3 that \( \vec{A} \) further satisfies the syntactic requirements for an attacker with recovery. We conclude in Theorem 2 by showing that \( \vec{A} \) also satisfies the semantic requirements for an attacker with recovery. (Indeed, \( \vec{A} \) is terminating.) This suffices to show that our solution is sound.

Subsequently, we prove that our solution is complete for terminating attackers, meaning that our solution returns a result \( \vec{A} \) precisely when a terminating TM-attacker exists. In order to prove that our solution is complete for terminating attackers, we first show that every terminating TM-attacker is discoverable using daisies with recovery, in Lemma 4.

Next, in Theorem 3 we use Lemma 4 to show that if a terminating attacker exists then the solution returns some result \( \vec{A} \), and if no terminating attacker exists then the solution states as much and terminates. This suffices to show that our solution is complete for the terminating subclass of attackers with recovery.

**Lemma 1 (Attacker Size).** Given a threat model \( \text{TM} = (P, (Q_i)_{i=0}^m, \phi) \), if the solution above returns a tuple \( \vec{A} \), it will have \( m+1 \) components.

**Proof.** If the solution above returns a tuple \( \vec{A} \), then it must have progressed to line 3. In this case, we know that \( R \) is non-empty in line 2. Then in line 5, the solution iterates \( m+1 \) times. In each iteration, the solution progresses through lines 5.1. through 5.9. Importantly, each of these lines terminates. (Line 5.1. terminates because \( r \models \text{F recover}_i \); for the other lines termination is obvious.) At line 5.9., the component \( A_i \) is appended to \( \vec{A} \). So after \( m+1 \) iterations of 5.1. through 5.9., there will be \( m+1 \) components \( A_0, \ldots, A_m \) in the tuple \( \vec{A} \), and we are done.

**Lemma 2 (Syntactic Requirements for an Attacker).** If the solution above yields a tuple \( (A_i)_{i=0}^m \), then each \( A_i \) is a deterministic process with no atomic propositions and the same inputs and outputs as the corresponding \( Q_i \).

**Proof.** First we will show that each \( A_i \) is deterministic, i.e., that each \( A_i \) satisfies determinism conditions (i) through (iv).

(i) Let \( T_i \) refer to the transitions of \( A_i \). Suppose \( T_i \) contains some transitions \( s \xrightarrow{l} a \) and \( s \xrightarrow{l} b \). If \( s \) is in recovery, then \( a \) and \( b \) must likewise be in recovery as there is no way to leave recovery. In this case, since the recovery process \( Q_i \) is assumed to be deterministic, it therefore follows that \( a = b \). What if \( s \) is not in recovery? Then the transition \( s \xrightarrow{l} a \) was either added at line 5.4. or line 5.5. For a contradiction assume \( a \neq b \).

- It cannot be the case that both transitions were added at line 5.4. because, at the end of line 5.4., the source state \( q_{0S} \) is incremented, violating the assumption that these distinct transitions begin at some state \( s \).

- If the former was added in 5.4. and the latter in 5.5., then the \( s \) must be an input-state. In this case, by construction the latter could not have the same label as the former, so we’ve arrived at a contradiction.

- If the latter was added in 5.4. and the former in 5.5. then the result is symmetric with the previous case.

- If both were added in 5.5. then they must have different labels, contradicting our assumption that they share the label \( l \).

So in all cases, assuming these are distinct transitions, we arrive at a contradiction. Therefore they must not be distinct transitions, so \( a = b \). We conclude that the transition relation \( T_i \) of \( A_i \) can be expressed as a (possibly partial) function \( S \times (I \cup O) \rightarrow S \).

(ii) Let \( s \) be a non-deadlock state in \( A_i \). If \( s \) is in recovery then, as the recovery is deterministic, \( s \) is either an input state or an output state, but not both. Else \( s \) is in the attack component. Consider this case. Clearly at line 5.4., some transition from \( s \) to either the next attack state or recovery was added, in order to simulate a corresponding transition taken by the daisy with recovery in the violating run \( r \). If \( s \) is an input state, then it has one outgoing input transition added at line 5.4., and potentially more outgoing input transitions added at line 5.5., but no other outgoing transitions. If \( s \) is an output state, then it has one outgoing output transition added at line 5.4., and no other outgoing transitions. Either way, \( s \) is either an input state or an output state, but not both.
Lemma 3 (Syntactic Requirements for an Attacker with Recovery). If the solution above yields a tuple $(A_i)_{i=0}^{m}$, then each $A_i$ consists of a finite DAG ending at the recovery $Q_i$.

Proof. Suppose the solution above yields a tuple $\vec{A} = (A_i)_{i=0}^{m}$. Choose a natural number $0 \leq i \leq m$ arbitrarily. We want to show that $A_i$ consists of a finite DAG ending at the recovery $Q_i$. Transitions outside of recovery in $A_i$ are added in lines 5.4. and 5.5. The number of such transitions added is linear in the product of the size of the interface $Q_i$ and the length of the finite prefix $\text{PRECOVERY}_i$. Hence, the part of $A_i$ outside recovery is clearly finite. But is it a DAG? The transitions added by line 5.4. form a line from $a_0, a_1, \ldots, a_{d-1}, q_0$. This line is totally ordered by the incrementing subscript on the state name, and is bounded in length by the size of $\text{PRECOVERY}_i$, and ends in recovery because of line 5.7. Every other pre-recovery transition is added by line 5.5. and points to $q_0$. So the pre-recovery portion is a DAG ending in $q_0$, the initial state of $Q_i$. As $0 \leq i \leq m$ was arbitrary, it follows that the pre-recovery portion of each $A_i$ is a DAG ending in the corresponding $q_0'$; so the lemma holds and we are done.

Theorem 2. The solution above is sound.

Proof. Suppose given some threat model $TM = (P, (Q_i)_{i=0}^{m}, \phi)$, the solution above returns $\vec{A}$. For convenience, define $\text{ATTACKEDSYSTEM}$ and $\text{DAISYSYSTEM}$ as we did previously. We need to show that $\vec{A}$ is, in fact, a TM-attacker with recovery. By Lemma 2, $\vec{A}$ is a tuple $(A_i)_{i=0}^{m}$ having $m + 1$ components. By Lemma 1 each component $A_i$ is a deterministic process with no atomic propositions and the same inputs and outputs as the corresponding $Q_i$. And by Lemma 3 each component $A_i$ consists of a finite DAG ending at the initial state of the corresponding $Q_i$. All that remains is to show that $\text{ATTACKEDSYSTEM}$ has at least one run violating $\phi$.

Let $r = s_0 \xrightarrow{l_0} s_1 \xrightarrow{l_1} \ldots$ be the run of $\text{DAISYSYSTEM}$ which was selected in line 3. of our solution. (Presumably we made it to line 3., since the solution returned an attacker $\vec{A}$.) We will show that $\text{ATTACKEDSYSTEM}$ can simulate the run $r$ with some run $r^{(m)}$ that violates $\phi$. We will make an inductive argument, albeit, one that terminates after $m + 1$ steps. Let $\zeta(i)$ be the inductive claim:

$$\zeta(i) := \exists r^{(i)} \in \text{runs}\left(P \parallel A_0 \parallel \cdots \parallel A_i \parallel R\text{DAISY}(Q_{i+1}) \parallel \cdots \parallel R\text{DAISY}(Q_m)\right)$$

such that $r^{(i)} \not\models \phi$ and in $r^{(i)}$, all the $A_i$s eventually recover.

In the base case, $i = -1$, and the system is $\text{DAISYSYSTEM}$. Setting $r^{(-1)} = r$, we see that $\text{DAISYSYSTEM}$ satisfies these conditions, thus $\zeta(-1) = \text{true}$. We now move on to the inductive step.

Assume $\zeta(i)$ holds. If $i = m$, then the entire theorem holds and we are done. Assume $i < m$ and let $r^{(i)}$ be the run discovered in the previous iteration of this argument.

Case 1: Suppose that in $r^{(i)}$, the process $R\text{DAISY}(Q_{i+1})$ begins in the recovery state $a_0^{i+1}$. Since the daisy portion of $R\text{DAISY}(Q_{i+1})$ is unreachable from the recovery portion $Q_{i+1}$, and since $Q_{i+1}$ uses all of its inputs and outputs, it

(iii) Every input state in recovery is input-enabled because we assume the recovery is deterministic. Every input state before recovery is input-enabled because of line 5.5.

(iv) Each output state in recovery has only one outgoing transition because we assume recovery is deterministic. Each output state before recovery has only one outgoing transition, as transitions are only added (once) in line 5.4. or (never, if the state is an output state) in line 5.5. So, every output state in $A_i$ has only one outgoing transition.
immediately follows that we could delete the daisy portion of RDAISY(Q_{t+1}) entirely and still find the same run r^{(i)}.

For convenience, let Q_{t+1} denote the result of this deletion, i.e., Q_{t+1} is the same as Q_{t+1} except that its initial state is additionally labeled recover_{t+1}. Define the following three systems:

\[
\begin{align*}
S1 &= P \parallel A_0 \parallel \cdots \parallel A_i \parallel Q_{t+1}' \parallel \text{RDAISY}(Q_{t+2}) \parallel \cdots \parallel \text{RDAISY}(Q_m) \\
S2 &= P \parallel A_0 \parallel \cdots \parallel A_i \parallel Q_{t+1} \parallel \text{RDAISY}(Q_{t+2}) \parallel \cdots \parallel \text{RDAISY}(Q_m) \\
S3 &= P \parallel A_0 \parallel \cdots \parallel A_i \parallel A_{i+1} \parallel \text{RDAISY}(Q_{t+2}) \parallel \cdots \parallel \text{RDAISY}(Q_m)
\end{align*}
\]

So modulo the propositions recover_0 through recover_{t+1}, r^{(i)} is also a run of S1, and thus also S2. But in this case, the solution will return A_{i+1} = Q_{t+1}, so it follows that r^{(i)} is (again, modulo the recover propositions) also a run of S3. Call this run r^{(i+1)}.

**Case 2:** Suppose that in r^{(i)}, the process RDAISY(Q_{t+1}) begins in the initial daisy state d_0^{t+1}. Notice that other than Case 1, this is the only remaining possibility, since the initial states of RDAISY(Q_{t+1}) are q_0^{t+1} and d_0^{t+1}. Let t_0, ..., t_j be the sequence of transitions that RDAISY(Q_{t+1}) takes from d_0^{t+1} in the prefix PRECOVERY_{t+1} defined at line 5.1.

Since both these claims hold, the augmented system can simulate the prefix PRECOVERY_{t+1}. But at the end of this prefix (and its simulation) RDAISY(Q_{t+1}) and A_{i+1} have both recovered, at which point they act identically. Keeping in mind that processes must use all their inputs and outputs, we conclude that the augmented system can simulate the entire run r^{(i)}. Call the simulating run of the augmented system r^{(i+1)}.

Since d_0^{t+1} and the attacker states a_i^{t+1} lack atomic propositions, it follows that modulo the propositions recover_0 through recover_{t+1}, r^{(i)} and r^{(i+1)} will induce identical traces. By construction, the propositions recover_0 through recover_{t+1} does not appear in the property φ, thus both of these traces violate φ. In both r^{(i)} and r^{(i+1)}, the processes P, A_0, ..., A_i, RDAISY(Q_{t+2}), ..., RDAISY(Q_m) progress through the same sequence of transitions and in the same order. Since we assumed these components recovered in r^{(i)} clearly they also recover in r^{(i+1)}. Can we say something similar with respect to RDAISY(Q_{i+1}) and A_{i+1}? Yes: In Case 1, the attacker component A_{i+1} begins in recovery, while in Case 2, Line 5.4. of our solution assures the transitions t_0 through t_j are consecutive, line 5.6. assures that t_j ends in recovery, and we just proved that transitions t_0 through t_j are taken by A_{i+1} in the run r^{(i+1)}. So in all cases, all the components recover in both runs. So the simulation requirements hold, i.e., φ \Rightarrow ψ.

Since we set φ_{i-1} = true, and proved that for −1 ≤ i < m, φ_{i} implies φ_{i+1}, we conclude by way of (finite) induction that A is a TM-attacker with recovery. This suffices to show our solution is sound, and we are done. □

**Lemma 4 (Terminating Attackers Induce Runs Discoverable by Daisies with Recovery).** Let TM = (P, (Q_i)_{i=0}^m, φ) be a threat model and A a terminating TM-attacker. Let ψ = (\bigwedge_{0 \leq i \leq m} F_{\text{recover}_i}) \Rightarrow φ. Then there exists a run r’ ∈ runs(P ∥ RDAISY(Q_0) ∥ ⋯ ∥ RDAISY(Q_m)) such that r’ \models ψ.

**Proof.** For convenience, define ATTACKEDSYSTEM and DAISYSYSTEM as follows:

\[
\begin{align*}
\text{ATTACKEDSYSTEM} &= P \parallel A_0 \parallel \cdots \parallel A_m \\
\text{DAISYSYSTEM} &= P \parallel \text{RDAISY}(Q_0) \parallel \cdots \parallel \text{RDAISY}(Q_m)
\end{align*}
\]
Since $\overline{A}$ is terminating, ATTACKEDSYSTEM has a run $r$ in which every $A_i$ visits the initial state $q_0^i$ of its recovery $Q_i$, and $r \not= \phi$. Let $r$ be some such run. We will show that DAISSYSTEM can "simulate" the run $r$, i.e., that DAISSYSTEM has some run $r'$ such that the only differences between $r$ and $r'$ are the names of the states the $A_i$s or RDAISY($Q_i$s) progress through prior to recovering.

First, consider $A_0$: let $r_0$ be the sequence of transitions that $A_0$ takes in the run $r$. Let $k_0 \geq 0$ be the smallest natural number such that in the $k_0$th step of $r_0$, $A_0$ is in the initial state $q_0^0$ of its recovery $Q_0$. If $k_0 = 0$, then $A_0$ begins in recovery; but once $A_0$ enters the recovery state $q_0^0$ it will remain in the states $S_{Q_0}$ of $Q_0$ forever. So in this case, $A_0$ simply acts like $Q_0$, and therefore $P \parallel$ DAISSYSTEM($Q_0$) $\parallel A_1 \parallel \cdots \parallel A_m$ can simulate the run $r$. On the other hand, what if $k_0 > 0$? Then the first $k_0 - 1$ transitions in $r_0$ are taken within the finite DAG part of $A_0$: the $k_0$th transition goes from the finite DAG portion to the recovery $q_0^0$; and all the subsequent transitions take place in $Q_0$ beginning at its initial state $q_0^0$. The first $k_0 - 1$ transitions can all be simulated via self-loops in the daisy part of the RDAISY($Q_0$), e.g., $a_0^0 \xrightarrow{x^0_i} a_0^0 \xrightarrow{x^0_j} a_0^0$. The $k_0$th transition can be simulated via a transition into recovery, e.g., $a_0^0 \xrightarrow{y^0_k} q_0^0$. Every subsequent transition will be the same, taking place in the recovery $Q_0$. This suffices to show that $P \parallel$ DAISSYSTEM($Q_0$) $\parallel A_1 \parallel \cdots \parallel A_m$ can simulate the run $r$. By applying this same argument inductively on the remaining $A_i$s, we find that DAISSYSTEM can simulate $r$.

Let $r'$ be a run of DAISSYSTEM simulating $r$. Then for each $0 \leq i \leq m$, by construction, there exists some $k_i \geq 0$ such that DAISSYSTEM($Q_i$) recovers at step $k_i$ of $r_i$. Therefore $r' \models \bigwedge_{0 \leq i \leq m} F_{\text{recover}_i}$. Moreover, $P$ progresses through the same sequence of states in $r'$ as it does in $r$. Since the $\text{recover}_i$ propositions do not occur in the property $\psi$, and the daises with recovery do not have any other atomic propositions, and the trace induced by $r$ violates $\psi$, it therefore follows that the trace induced by $r'$ likewise violates $\psi$. Therefore $r'$ violates $\psi$. We conclude that $r'$ violates $\psi$, hence DAISSYSTEM $\not\models \psi$ and we are done.

**Theorem 3.** The solution above is complete, for terminating attackers.

**Proof.** First we show that the solution terminates on all inputs. We assume the model checker terminates, so we always progress to line 2. If $R = \emptyset$ then we terminate at line 2. If $R \not= \emptyset$, then we progress to line 3. Next we loop over $0 \leq i \leq m$. Each step of the loop terminates in runtime bounded by the length of the shortest prefix of the violating run ending in recovery for the corresponding attacker component. Since the run is assumed to satisfy $F_{\text{recover}_i}$ for each $0 \leq i \leq m$, this shortest prefix must exist. Thus the loop completes in finite time. Once the loop completes we return the tuple $\overline{A}$ and we are done. We conclude that the solution terminates on all inputs.

But, does the solution at least attempt to return an attacker, whenever a terminating attacker exists? And when no terminating attacker exists, does it state as much? We will answer both questions affirmatively, beginning with the first.

Suppose a terminating attacker exists. In Lemma\textsuperscript{[18]} we proved that terminating attackers induce runs discoverable by daises with recovery. It follows that at line 2 of our solution the set $R$ will be non-empty. Thus the solution will at least attempt to return an attacker, i.e., it will progress past line 3.

Now we address the second question. Suppose no terminating attacker exists. Then the set $R$ will be empty, and the solution will terminate at line 2. As both questions are answered in the affirmative, we conclude that the solution is complete for terminating attackers.

## 5 Case Study: TCP

Below we first describe our implementation then the details of our case study (TCP).

**Implementation** We implemented our solutions in an open-source tool called KORG\textsuperscript{[4]}. We say an attacker $\overline{A}$ for a threat model TM = ($P$, ($Q_i$)$_{i=0}^m$, $\phi$) is a centralized attacker if $m = 0$, or a distributed attacker, otherwise. In other words, a centralized attacker has only one attacker component $\overline{A} = (A)$, whereas a distributed attacker has many attacker components $\overline{A} = (A_i)_{i=0}^m$. KORG handles $\exists\text{ASP}$ and $\forall\exists\text{ASP}$ for liveness and safety properties for a centralized attacker. KORG is implemented in PYTHON 3 and uses the model-checker SPIN\textsuperscript{[18]} as its underlying verification engine.

\textsuperscript{[4]} Named after the Korg microKORG synthesizer, with its dedicated “attack” control on Knob 3. Code and models are freely and openly available at [https://github.com/maxvonhipelpe/AttackerSynthesis](https://github.com/maxvonhipelpe/AttackerSynthesis)

16
| Attacker: | Centralized | Distributed |
|----------|-------------|-------------|
| Recovery? | ∃ Problem   | ∃ Problem   |
| With     | ✓            | ✗           |
| Without  | ✓            | ✗           |

Table 1: ✓ or ✗ denote if the solution to a problem is or is not implemented in KORG, for both liveness and safety properties.

TCP is a fundamental Internet protocol consisting of three stages: connection establishment, data transfer, and connection tear-down. We focus on the first and third stages, which jointly we call the connection routine. Our approach and model (see Fig. 8) are inspired by SNAKE [22]. Run-times and results are listed in Table 2.

![TCP threat model block diagram](image)

**Threat Models** We use channels to build asynchronous communication out of direct (rendezvous) communication. Rather than communicating directly with the NETWORK, the peers communicate with the channels, and the channels communicate with the NETWORK, allowing us to model the fact that packets are not instantaneously transferred in the wild. We use the shorthand CHAN!MSG to denote the event where MSG is sent over a channel CHAN; it is contextually clear who sent or received the message. TCP exists in the Transport Layer of the internet, an upper layer reliant on the lower Link Layer and Internet Layer. We abstract the lower network stack layer TCP relies on with NETWORK, which passes messages between 1TON ∥ 2TON and NT01 ∥ NT02. We model the peers symmetrically.

Given a property ϕ about TCP, we can formulate a threat model TM as follows, where we assume the adversary can exploit the lower layers of a network and ask if the adversary can induce TCP to violate ϕ:

$$TM = (\text{Peer 1} ∥ \text{Peer 2} ∥ 1TON ∥ 2TON ∥ NT01 ∥ NT02, (\text{NETWORK}), \phi)$$

We consider the properties ϕ₁, ϕ₂, ϕ₃, giving rise to the threat models TM₁, TM₂, TM₃.

**TM₁: No Half-Closed Connection Establishment** The safety property ϕ₁ says that if Peer 1 is in Closed state, then Peer 2 cannot be in Established state.

$$ϕ₁ = G(\text{Closed}_1 \implies \neg \text{Established}_2)$$

KORG discovers an attacker that spoofs the active participant in an active-passive connection establishment (see message sequence chart in Fig. 10), as described in [15]. Note that our model does not capture message sequence numbers and in the real world the attacker also needs to guess the sequence number of the passive peer.
Fig. 9: A TCP peer. For \( i = 1, 2 \), if this is PEER \( i \), then \( \text{SND} := \text{ITON} \) and \( \text{RCV} := \text{NTOI} \). All the states except \( i_0, \ldots, i_5 \), and \( \text{End} \) are from the finite state machine in the TCP RFC [36]. The RFC diagram omits the implicit states \( i_0, \ldots, i_5 \), instead combining send and receive events on individual transitions. In the RFC, Closed is called a “fictional state”, where no TCP exists. We add a state \( \text{End} \) to capture the difference between a machine that elects not to instantiate a peer and a machine that is turned off. We label each state \( s \) with a single atomic proposition \( s_i \). Dashed transitions are timeout transitions, meaning they are taken when the rest of the system deadlocks.

Fig. 10: Time progresses from top to bottom. Labeled arrows denote message exchanges over implicit channels. The property is violated in the final row; after this recovery may begin. Note that the initial \( \text{ACK} \) injected to Peer 1 is unnecessary. Because SPIN attempts to find counterexamples as quickly as possible, the counterexamples it produces are not in general minimal.

**TM2: Passive-Active Connection Establishment Eventually Succeeds** The liveness property \( \phi_2 \) says that if it is infinitely often true that PEER 1 is in Listen state while PEER 2 is in SYN Sent state, then it must eventually be true that PEER 1 is in Established state.

\[
\phi_2 = (\text{GF}(\text{Listen}_1 \land \text{SYN Sent}_2)) \implies \text{F Established}_1
\]  

(10)

KORG discovers an attack where an \( \text{ACK} \) packet is injected to each peer, and a \( \text{SYN} \) packet from PEER 2 is dropped. The corresponding attacker code is given in the PROMELA language of SPIN in Fig. [11] The attacker in Fig. [11] also induces deleterious behavior not captured by violation of \( \phi_2 \), where the system deadlocks in (Listen, SYN Sent).
\begin{align*}
\text{TM}_3: \text{Peers Do Not Get Stuck} & \quad \text{The safety property } \phi_3 \text{ says that the two peers will never simultaneously deadlock outside their End states. Let } S_i \text{ denote the set of states in Fig. 9 for } \text{Peer } i, \text{ and } S'_i = S_i \setminus \{\text{End}\}.
\phi_3 = \bigwedge_{s_1 \in S'_1} \bigwedge_{s_2 \in S'_2} \neg \text{FG}(s_1 \land s_2) \quad (11)
\end{align*}

For the problem with recovery, KORG discovers an attacker that injects a SYN packet to each peer, causing both to transition into SYN Received, where they get stuck (violating \( \phi_3 \)).

\textbf{Performance} \quad \text{Performance results for Case Study are given in Table 2. The discovered attackers either implement known attacks or reproduce known bugs in real TCP implementations. Our success criteria was to produce realistic attackers faster than an expert human could with pen-and-paper. We discovered attackers in seconds or minutes as shown in Table 2.}

\begin{table}[h]
\centering
\begin{tabular}{|c|c|c|c|}
\hline
Property & Avg. Runtime (s) & Unique Attacker & Unique Attackers Found \\
\hline
\exists ASP & R-\exists ASP & \exists ASP & R-\exists ASP \\
\hline
\phi_1 & 0.43 & 3.39 & 7 & 1 \\
\phi_2 & 0.75 & 3.29 & 4 & 1 \\
\phi_3 & 122.88 & 2469.74 & 3 & 1 \\
\hline
\end{tabular}
\caption{Since we use size-1 channels, we allow KORG to configure SPIN with partial order reduction turned on, which improves overall runtime. For each property \( \phi_i \), we asked KORG to generate ten attackers with recovery, and ten without, using a 16Gb 2020 quad-core Intel© Core™ i7-1185G7 CPU running Linux Mint 20.1 Cinnamon. KORG may generate duplicate attackers, so for each property (Column 1), we list the time taken to generate a unique attacker without recovery (Column 2) or with recovery (Column 3), and the total number of unique attackers found without recovery (Column 4) or with recovery (Column 5). For example, for \( \phi_3 \), out of 10 attackers with recovery generated over the course of about 41 minutes, one was unique and the rest were duplicates. Instructions and code to reproduce these results are given in the GitHub repository.}
\end{table}

We chose TCP connection establishment for our case study because it is simple and well-understood. Across three properties (two safety and one liveness), with and without recovery, KORG synthesized attackers exhibiting attack strategies that have worked or could work against some real-world TCP implementations, modulo sequence numbers. The synthesized attackers are simple, consisting of only a few lines of code, but our TCP model is also simple since we omitted sequence numbers, congestion control, and other details. Moving forward, we want to apply KORG to more complicated models and discover novel exploits.

\section{Related Work}

Prior works formalized security problems using game theory (e.g., FLIPT [41], [26]), “weird machines” [9], attack trees [45], Markov models [40], and other methods. Prior notions of attacker quality include \( O \)-complexity [8], expected information loss [38], or success probability [32][44], which is similar to our concept of \( \forall \) versus \( \exists \)-attackers. The formalism of [44] also captures attack consequence (cost to a stakeholder).

Nondeterminism abstracts probability, e.g., a \( \forall \)-attacker is an attacker with \( P(\text{success}) = 1 \), and, under fairness conditions, an \( \exists \)-attacker is an attacker with \( 0 < P(\text{success}) < 1 \). Probabilistic approaches are advantageous when the
existence of an event is less interesting than its likelihood. For example, a lucky adversary could randomly guess my RSA modulus, but this attack is too unlikely to be interesting. We chose to use nondeterminism over probabilities for two reasons: first, because nondeterministic models do not require prior knowledge of event probabilities, but probabilistic models do; and second, because the non-deterministic model-checking problem is cheaper than its probabilistic cousin [42]. Nevertheless, we believe our approach could be extended to probabilistic models in future work. Katoen provides a nice survey of probabilistic model checking [24].

Attacker synthesis work exists in cyber-physical systems [34, 32, 28, 29, 32], most of which define attacker success using bad states (e.g., reactor meltdown, vehicle collision, etc.) or information theory (e.g., information leakage metrics). Problems include the actuator attacker synthesis problem [29]; the hardware-aware attacker synthesis problem [39]; and the fault-attacker synthesis problem [5]. Defensive synthesis also exists [2].

Maybe the most similar work to our own is PROVERIF [7], which verifies properties of, and generates attacks against, cryptographic protocols. We formalize the problem with operational semantics (processes) and reduce it to model checking, whereas PROVERIF uses axiomatic semantics (PROLOG clauses) and reduces it to automated proving. Another similar tool is NETSMC [47], a model-checker that efficiently finds counter-examples to security properties of stateful networks. NETSMC is a model-checker, whereas our approach is built on top of a model checker, so in theory our approach could be implemented on top of NETSMC.

Existing techniques for automated attack discovery include model-guided search [22, 19] (including using inference [10]), open-source-intelligence [46], bug analysis [20], and genetic programming [25]. The generation of a failing test-case for a protocol property is not unlike attack discovery, so [31] is also related.

TCP was previously formally studied using a process language called SPEX [37], Petri nets [16], the HOL proof assistant [6], and various other algebras (see Table 2.2 in [30]). Our model is neither the most detailed nor the most comprehensive, but it captures all possible establishment and tear-down routines, and is tailored to our framework.

This paper focuses on attacker synthesis at the protocol level, and thus differs from the work reported in [23] in two ways: first, the work in [23] synthesizes mappings between high-level protocol models and execution platform models, thereby focusing on linking protocol design and implementation; second, the work in [23] synthesizes correct (secure) mappings, whereas we are interested in synthesizing attackers.

7 Conclusions and Future Work

We present a novel formal framework for automated attacker synthesis. The framework includes an explicit definition of threat models and four novel, to our knowledge, abstract categories of attackers. We formulate four attacker synthesis problems, and propose solutions to two of them by program transformations and reduction to model-checking. We prove our solutions are sound. We prove our solution to the first problem is complete generally, and that our solution to the second problem is complete over a restricted class of attackers. Proofs are provided in Section 4. Finally, we implement our solutions for the case of a centralized attacker in an open-source tool called KORG, apply KORG to the study of the TCP connection routine, and discuss the results. KORG and the TCP case study are freely and openly available, so our results are easy to reproduce.

8 Erratum

In a previous version of this paper [17], we proposed an incorrect solution to the R-∀ASP, which employed a model-checker to find a violating run \( r \) described as a lasso \( \alpha \cdot \beta^\omega \), then let each attacker component \( A_i \) consist of a process reproducing the component of \( \alpha \) in the interface of \( Q_i \), followed by the recovery \( Q_i \). That solution was flawed because it assumed that at the end of \( \alpha \) (no later and no sooner), all of the daisies with recovery would recover. We illustrate how that assumption could fail below.

Consider the processes \( Q \) and \( P \) in Fig. 12 and let \( TM = (P, (Q), \phi) \) be a threat model where \( \phi = G\neg \text{dead} \). A lasso run \( r = \alpha \cdot \beta^\omega \) of \( P \parallel \text{DAISY}(Q) \) violating \( \psi = \text{Frecover} \implies (G\neg \text{dead}) \) is given in Fig. 13. Since at least one such run exists, in both our prior (incorrect) and revised (correct) solutions, \( R \neq \emptyset \).
Our prior (incorrect) solution to the $\exists$ASP would produce the following process, $A$.

But critically, $P \parallel A$ has no runs, hence $A$ is not actually a TM-attacker. The problem is that we created the attacker program by projecting the prefix $\alpha$ onto the interface of $Q$, and gluing the result directly to the recovery $Q$. Instead, we should have projected the prefix of $r$ up until recover, and glued that result to recovery. This is exactly what we do in our updated solution, in this case, yielding the terminating attacker $A'$ illustrated in Fig. 15. The attacked system $P \parallel A'$ yields the run

$$r' = (p_0, a_0') \xrightarrow{\text{TRIGGER!}} (p_1, q_0) \xrightarrow{\text{BAD!}} (p_0, q_1) \xrightarrow{\text{DOBAD!}} (p_4, q_1) \xrightarrow{\text{DOBAD!}} (p_4, q_1) \xrightarrow{\text{DOBAD!}} \ldots$$

which violates $\psi = (F \text{recover}) \Rightarrow (G \neg \text{dead})$. 

---

Fig. 12: Left: the process $Q$. Notice that $Q$ is deterministic. Right: the process $P$. Observe that $P \parallel Q \models G \neg \text{dead}$.

Fig. 13: An example (lasso) run $r$ of $P \parallel \text{RDAISY}(Q)$ violating $\psi = (F \text{recover}) \Rightarrow (G \neg \text{dead})$.

Fig. 14: The “attacker” component produced by our prior (incorrect) solution given the example threat model.
Fig. 15: The corrected attacker component $A'$ produced by our updated solution given the example threat model TM.

Acknowledgments This material is based upon work supported by the National Science Foundation under NSF SaTC award CNS-1801546. Any opinions, findings, and conclusions or recommendations expressed in this material are those of the author(s) and do not necessarily reflect the views of the National Science Foundation. The authors thank four anonymous reviewers. Additionally, the first author thanks Benjamin Quiring, Dr. Ming Li, and Dr. Frank von Hippel.

References

1. Alur, R., Tripakis, S.: Automatic synthesis of distributed protocols. SIGACT News 48(1), 55–90 (2017)
2. Ardeshiricham, A., Takashima, Y., Gao, S., Kastner, R.: VeriSketch: Synthesizing secure hardware designs with timing-sensitive information flow properties. In: Proceedings of the 2019 ACM SIGSAC Conference on Computer and Communications Security. pp. 1623–1638 (2019)
3. Baier, C., Katoen, J.P.: Principles of model checking. MIT press (2008)
4. Bang, L., Rosner, N., Bultan, T.: Online synthesis of adaptive side-channel attacks based on noisy observations. In: 2018 IEEE European Symposium on Security and Privacy. pp. 307–322. IEEE (2018)
5. Barthe, G., Dupressoir, F., Fouque, P.A., Grégoire, B., Zapalowicz, J.C.: Synthesis of fault attacks on cryptographic implementations. In: Proceedings of the 2014 ACM SIGSAC Conference on Computer and Communications Security. pp. 1016–1027 (2014)
6. Bishop, S., Fairbairn, M., Mehnert, H., Norrish, M., Ridge, T., Sewell, P., Smith, M., Wansbrough, K.: Engineering with logic: Rigorous test-oracle specification and validation for TCP/IP and the sockets API. Journal of the ACM 66(1), 1–77 (2018)
7. Blanchet, B.: An Efficient Cryptographic Protocol Verifier Based on Prolog Rules. In: 14th IEEE Computer Security Foundations Workshop. pp. 82–96. IEEE Computer Society, Cape Breton, Nova Scotia, Canada (Jun 2001)
8. Branco, R., Hu, K., Kawakami, H., Sun, K.: A mathematical modeling of exploitations and mitigation techniques using set theory. In: 2018 IEEE Security and Privacy Workshops (SPW). pp. 323–328. IEEE (2018)
9. Bratus, S., Locasto, M.E., Patterson, M.L., Sassaman, L., Shubina, A.: Exploit programming: From buffer overflows to weird machines and theory of computation. USENIX: login 36(6) (2011)
10. Cho, C.Y., Babic, D., Poosankam, P., Chen, K.Z., Wu, E.X., Song, D.: MACE: Model-inference-assisted concolic exploration for protocol and vulnerability discovery. In: USENIX Security Symposium. vol. 139 (2011)
11. Chong, S., Gutman, J., Datta, A., Myers, A., Pierce, B., Schaumont, P., Sherwood, T., Zeldovich, N.: Report on the NSF workshop on formal methods for security. (2016)
12. Church, A.: Application of recursive arithmetic to the problem of circuit synthesis (7 1957), as cited in doi:10.2307/2271310.
13. Dijkstra, E.W., et al.: Notes on structured programming. http://www.cs.utexas.edu/users/EWD/EWD02xx/EWD249.PDF (1970), accessed: 11 May 2020
14. Duran, J.W., Ntafos, S.: A report on random testing. In: Proceedings of the 5th international conference on Software engineering. pp. 179–183. IEEE Press (1981)
15. Friedrichs, O.: A simple TCP spoofing attack. citi.umich.edu/u/provos/papers/secnet-spoof.txt (February 1997), accessed: 3 January 2020
16. Han, B., Billington, J.: Termination properties of TCP’s connection management procedures. In: International Conference on Application and Theory of Petri Nets. pp. 228–249. Springer (2005)
17. von Hippel, M., Vick, C., Tripakis, S., Nita-Rotaru, C.: Automated attacker synthesis for distributed protocols. In: International Conference on Computer Safety, Reliability, and Security. pp. 133–149. Springer (2020)
47. Yuan, Y., Moon, S.J., Uppal, S., Jia, L., Sekar, V.: NetSMC: A custom symbolic model checker for stateful network verification. In: 17th USENIX Symposium on Networked Systems Design and Implementation. USENIX Association, Santa Clara, CA (Feb 2020)