Automated Analysis of Scenario-based Specifications of Distributed Access Control Policies with Non-Mechanizable Activities *
(Extended Version)

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Abstract. The advance of web services technologies promises to have far-reaching effects on the Internet and enterprise networks allowing for greater accessibility of data. The security challenges presented by the web services approach are formidable. In particular, access control solutions should be revised to address new challenges, such as the need of using certificates for the identification of users and their attributes, human intervention in the creation or selection of the certificates, and (chains of) certificates for trust management. With all these features, it is not surprising that analyzing policies to guarantee that a sensitive resource can be accessed only by authorized users becomes very difficult. In this paper, we present an automated technique to analyze scenario-based specifications of access control policies in open and distributed systems. We illustrate our ideas on a case study arising in the e-government area.

1 Introduction

Access control aims at protecting data and resources against unauthorized disclosure and modifications while ensuring access to authorized users. An access control request consists of a subject asking to perform a certain action on an object of a system. A set of policies allows the system to decide whether access is granted or denied by evaluating some conditions on the attributes of the subject, the object, and the environment of the system (such as the identity, role, location, or time). For centralized systems, identifying subjects, objects, and the values of the attributes is easy since both subjects and objects can be adequately

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classified by identifiers that are assigned by the system itself. For open and distributed systems such as those based on web technology, the situation is more complex as web servers receive and process requests from remote parties that are difficult to identify and to bind with their attribute values. Hence, certificates or credentials, attesting not only the identity but also the attributes of parties, must be exchanged to correctly evaluate access control queries. In many situations, the creation and exchange of certificates require human intervention, e.g., to issue and sign a certificate or to pick one in a portfolio of available credentials. Furthermore—as observed in [18] among others—in distributed systems, a certificate can be accepted or rejected depending on the trust relation between the receiver and its issuer. Additional flexibility can be gained by chains of credentials and trust. In this context, guaranteeing that only trusted users can access sensitive resources becomes a daunting task.

1.1 Main Contributions

In this paper, we propose a technique for the automated analysis of access control systems (ACS) in presence of human activities for the creation and exchange of certificates together with trust management. Our approach combines a logic-based language with model checking based on Satisfiability Modulo Theories (SMT) solving. More precisely, we follow [17] and use Constraint Logic Programming (CLP) for the specification of policies and trust management with ideas adapted from [14]. The exchange of certificates and their interplay with the set of policies is modeled as a transition system of the type proposed in [19]. We show that interesting analysis problems of ACSs can be reduced to reachability problems. Our main contribution is a decidability result for the (bounded) reachability problem of a sub-class of transition systems that can encode the analysis of scenario-based specifications of ACSs, i.e. situations in which the exchange of certificates is constrained by a given causality relation. Another contribution is a technique to reduce the number of possible interleavings while visiting reachable states.

1.2 A Motivating Example: the Car Registration Office

We consider a simplified version of the Car Registration Office (CRO) application in [4]. It consists of a citizen wishing to register his new car via an on-line service provided by the CRO. An employee of the CRO, Ed, checks if the request can be accepted according to some criteria. If so, Ed must store the request in a central repository CRep, which, in turn, checks if Ed is entitled to do so. To be successful, the storage request must be supported by three certificates: ise saying that Ed is an employee of the CRO, ish saying that Helen is the head of the CRO and cans saying that Helen granted Ed the permission to store documents in CRep. Roles certificates must be signed by a trusted Certification Authority (CA) while Ed’s permission certificate is signed by Helen; if these were not the case, the certificates should be rejected because the principal that signed the properties is untrusted. The generation of certificates (depicted in Fig. 1) is
a non-mechanizable activity whose execution depends on decisions that are not modeled in the system but only on the human behavior. Another issue is how the certificates are sent to CRep in order to support Ed’s storage request. It can be Ed to send the certificates along with the request (user-pull) or it can be CRep to collect the necessary certificates upon reception of Ed’s request (server-pull).

**Organization of the paper**

In Section 2, we give an overview of the main features of our approach, which we then detail in Section 3 where we formalize a class of access control schemas. In Sections 4 and 5, we present our main contributions: an automated analysis technique of scenario-based specifications and a heuristics for its scalability. In Section 6, we conclude and discuss related work. Formal preliminaries, a complete derivation of the CRO main query and an implementation of the scenario by using DKAL language, can be found in the appendix Section 7.

2 Overview of the Main Features of our Approach

Our goal is to automatically analyze situations in which (a) certificates are created or exchanged because of human intervention, (b) there is a need to reason about chains of credentials to establish the validity of a certificate, and (c) message exchanges comply with a causality relation.

2.1 Certificates and non-mechanizable activities

Inspired by [17, 14], we use a variant of Constraint Logic Programming (CLP) to abstractly represent certificates as well as to specify and reason about the trust relationships among principals and the restrictions on delegating the ability to issue certificates.

**Example 1.** For the CRO scenario, the three certificates depicted in Fig. 1 can be expressed as the following CLP facts:

\[
\begin{align*}
(F1) & \text{ uknows(CA, a2i(Ed, ise))} \\
(F2) & \text{ uknows(CA, a2i(Helen, ish))} \\
(F3) & \text{ uknows(Helen, a2i(Ed, cans))},
\end{align*}
\]
where \( \text{uknows} \) represents the knowledge of a principal resulting from non-mechanizable activities only, called \emph{internal} knowledge, and \( \text{a2i} \) is a constructor for the piece of knowledge about the binding of a property (e.g., being an employee, \( \text{ise} \)) with a principal (e.g., \( \text{Ed} \)).

### 2.2 Exchange of certificates among principals

Distributed access control is based on exchanging certificates among principals so that access decisions can be taken by one principal with all the necessary information. So, we need to specify the actions that change the state of the system, that is the content of the network and the internal knowledge of the principals involved. To this end, we use the notion of transition system introduced in [19] for access control systems as follows. The network of messages is modeled by a ternary predicate \( \text{msg} \) with three arguments: the sender, the payload, and the receiver of the message. The action of \( p \) sending a message with payload \( x \) to \( q \) can be written as a transition

\[
\text{knows}(p, x) \Rightarrow \oplus \text{msg}(p, \text{said}(x), q)
\]

where \( \text{knows} \) represents the knowledge of a principal, both internal and acquired from the reception of messages from other principals, and \( \text{said} \) transforms a piece of knowledge into an assertion that can be communicated to other principals. The fact that internal knowledge is knowledge can be expressed by the CLP rule

\[
\text{knows}(p, x) \leftarrow \text{uknows}(p, x)
\]

and the action of \( q \) receiving a message from \( p \) with \( s \) as payload is written as

\[
\text{knows}(q, s2i(p, s)) \leftarrow \text{msg}(p, s, q)
\]

where \( s2i \) is a constructor for the piece of knowledge about the binding of the utterance \( s \) with a principal \( p \).

\textbf{Example 2.} For example, the action of \( \text{CA} \) sending the certificate that \( \text{Ed} \) is an employee to \( \text{Ed} \) himself can be formalized as an instance of (1), the reception of such a certificate by \( \text{Ed} \) as an instance of (3), and the derivation that \( \text{Ed} \) knows that \( \text{CA} \) has uttered (and signed) the property about \( \text{Ed} \) being an employee—formally, \( \text{knows}([\text{Ed}, s2i(\text{CA}, \text{said(a2i(Ed, ise))})]) \)—as an application of fact (F1) and rule (2). Notice that \( \text{Ed} \) cannot claim to know that he is an employee since he does not know whether \( \text{CA} \) is trusted on emitting this type of utterances. For this, suitable trust relationships should be specified.

### 2.3 Trust relationships among principals

We use again CLP rules. One rule is generic while the others are application dependent. The generic rule is

\[
\text{knows}(p, x) \leftarrow \text{knows}(p, s2i(q, \text{said}(x))) \land \text{knows}(p, a2i(q, \text{tdOn}(x)))
\]
saying that a principal \( p \) may expand its knowledge to include the piece of information \( x \) as soon as another principal \( q \) has uttered \( \text{said}(x) \) and \( q \) is trusted on the same piece of knowledge \( x \) (the last part is encoded by the term \( a2i(q, \text{tdOn}(x)) \)).

Example 3. In the case of the CRO, we need also to consider the following four specific CLP rules, that encode the trust relationships among the various principals:

\[
\begin{align*}
(P1) \quad & \text{knows} (\text{CRep}, a2i(p, \text{cans})) \leftarrow \text{knows} (\text{CRep}, a2i(q, \text{ish})) \\
& \quad \land \text{knows} (\text{CRep}, a2i(p, \text{ise})) \land \text{knows} (\text{CRep}, s2i(q, \text{said}(a2i(p, \text{cans})))) \\
(P2) \quad & \text{knows}(p, a2i(\text{CA}, \text{tdOn}(x))) \\
(P3) \quad & \text{knows}(p, a2i(q, \text{tdOn}(s2i(\text{CA}, \text{said}(x))))) \\
(P4) \quad & \text{knows}(p, a2i(q, \text{tdOn}(s2i(r, \text{said}(a2i(q, \text{cans})))))) \leftarrow \text{knows}(p, a2i(r, \text{ish})),
\end{align*}
\]

\( (P1) \) says that a principal \( p \) can store documents in the \text{CRep} if he is an employee of the CRO and his head permits it, \( (P2) \) says that the content of any utterance of the \text{CA} is trusted, \( (P3) \) says that an utterance of a principal repeating an utterance of the \text{CA} is trusted, and finally \( (P4) \) says that the head of the CRO is trusted when emitting an utterance granting permission to store documents in the \text{CRep} to a principal.

\( \Box \)

2.4 Automated analysis of scenarios

The formal framework sketched above allows us to develop automated analysis techniques to verify the availability (policies suitable for scenario’s execution) or the security (critical operations performed by trusted principals) of typical scenarios in which an ACS should operate. Availability implies that the policies are not too restrictive to prevent the scenario to be executable while security means that only trusted principals are granted access to sensitive resources or perform sensitive operations. Both problems can be reduced to check whether, after performing a sequence of non-mechanizable activities and exchanging messages among principals, it is possible to reach a configuration of the network in which an access control query (e.g., in the CRO, “Can Ed store the citizen’s request in \text{CRep}?”) gets a positive or a negative answer.

In other words, we want to solve problems as stated by the following definition:

Definition 1 (Reachability problems). Given the following conditions:

- let the network be initially empty (formally, \text{msg} is interpreted as an empty relation),
- \( H_0 \) be a set of facts derived from non-mechanizable activities (e.g., \( (F1) \), \( (F2) \) and \( (F3) \) described in Example 1),
- and \( G \) be a conjunction of \text{knows}-facts describing an access control query (e.g., \text{knows}(\text{CRep}, a2i(\text{Ed, cans})) for the CRO example)
we aim to check if does there exist a sequence of \( n \) instances of the transition rule (1) and a sequence \( H_1, \ldots, H_n \) of knows-facts derived from non-mechanizable activities, such that \( G \) is satisfied in the final state?

To practically answer this question, initially we need to compute the fix-point of the facts in \( H_0 \) with the CLP rules (2), (3), (4) and those formalizing specific trust relations. This process must be repeated at each step \( i = 1, \ldots, n \) with the facts describing the content of the network (derived by applying (3) at step \( i-1 \)), those in the set \( H_i \), and the CLP rules. Since more than one transition (1) can be enabled at any given step \( i \), it is necessary, in general, to consider several possible execution paths.

Not surprisingly, the reachability problem turns out to be quite difficult. Fortunately, in scenarios with constrained message exchanging (e.g., the user-pull or the server-pull configurations considered for the CRO above), the reachability problem becomes simpler. It is possible to fix a bound \( n \) of transitions to consider and apply a reduction technique to decrease the number of different execution paths to be explored as we will see in Sections 4 and 5.

3 A Class of Access Control Schemas

According to [19], we report, in the following, the definition of access control schema (in short \( \mathcal{ACS} \)).

**Definition 2 (Access Control Schema).** An \( \mathcal{ACS} \) is a transition system

\[
(S, Q, \vdash, \Psi),
\]

where \( S \) is a set of states, \( Q \) is a set of queries, \( \Psi \) is a set of state-change rules, and \( \vdash \subseteq S \times Q \) is the relation establishing if a query \( q \in Q \) is satisfied in a given state \( \gamma \in S \), written as \( \gamma \vdash q \).

For \( s, s' \in S \) and \( \psi \in \Psi \), we write \( s \rightarrow_{\psi} s' \) when the change from \( s \) to \( s' \) is allowed by \( \psi \). The reflexive and transitive closure of \( \rightarrow_{\psi} \) is denoted by \( \rightarrow_{\psi}^{\ast} \).

Given an \( \mathcal{ACS} \) \( (S, Q, \vdash, \Psi) \), an instance \( (s, q, \psi) \) of the reachability problem (see Definition 1) (where \( s \in S \), \( q \in Q \), and \( \psi \in \Psi \)) consists of asking whether there exists an \( s' \in S \) such that \( s \rightarrow_{\psi}^{\ast} s' \) and \( s' \vdash q \).

3.1 The substrate theory \( T_5 \)

We define a class of \( \mathcal{ACSs} \) by using formulae of (many-sorted) first-order logic [12] to represent states and transitions. To do this formally, we need to introduce a substrate theory \( T_5 \), i.e., a set of formulae that abstractly specifies the basic data-structures and operations relevant for both access control and trust management. The theory contains a (countably) infinite set of constants of sort \( \text{Principal} \) to identify users, suitable operations to build \( \text{Attribute} \) values, and the functions \( a2i : \text{Principal} \times \text{Attribute} \rightarrow \text{Infon} \), \( s2i : \text{Principal} \times \text{Speech} \rightarrow \text{Infon} \), \( \text{said} : \)}
Infon → Speech, tdOn : Infon → Attribute, that have been already informally described in Section 2.

Moreover the substrate theory $T_S$ contains the predicate symbol prim : Attribute that intuitively characterizes the set of “primitive” attributes, i.e., those already in the substrate that are not created by the “function” tdOn (e.g., ise, ish, cans for the CRO example). So, it is necessary to add to the substrate theory the following axiom

$$\forall x, a. \text{tdOn}(x) = a \Rightarrow \neg \text{prim}(a)$$

where $x$ is a variable of sort Infon and $a$ is a variable of sort Attribute.\(^1\) Another important aspect we want to remark is that, even if in this paper we assume, for the sake of simplicity, the standard situation (see, e.g., [20]) where insecure communication channels between each pair of principals are always available, it is easy to extend the substrate theory by adding axioms to characterize the “topology” of the system.

We recall that the theory $T_S$ identifies a class of structures that are models of all formulae in $T_S$ and say that a formula $\varphi$ is satisfiable modulo $T_S$ iff there exists a model of $T_S$ that makes $\varphi$ true.

### 3.2 The set $S$ of states

We consider the two predicate symbols uknows : Principal × Infon and msg : Principal × Speech × Principal already introduced in Section 2. We assume the availability of a finite set $Po$ of CLP rules, also called policies, of the form

$$A_0(\bar{x}) \leftarrow A_1(\bar{x}, \bar{y}) \wedge \cdots \wedge A_n(\bar{x}, \bar{y}) \wedge \xi(\bar{x}, \bar{y}),$$

where $\bar{x}$ and $\bar{y}$ are tuples of variables, $A_0$ is knows, $A_i \in \{\text{knows}, \text{uknows}\}$ for $i = 1, \ldots, n$, and $\xi$ is a quantifier-free formula of the substrate theory $T_S$. We assume $Po$ to always contain (2), (3), and (4). Given a set $F$ of constrained ground facts and the set $Po$ of policies, the set $S$ of states contains all the constrained ground facts obtained by computing the least-fixpoint $\text{lfp}(F \cup Po)$ of the ground immediate consequence operator on $F \cup Po$ (see, e.g., [17]).

### 3.3 The set $Q$ of queries and the satisfaction relation $\vdash$

A query is a conjunction of ground facts of the form $\text{uknows}(p, x) \leftarrow \xi(p, x)$. We define $\vdash$ to be the standard consequence relation $\models$ of first-order logic [12].

\(^1\) Indeed, the models of the theory considered here are a super-class of those considered in [14]. Here, we trade precision for the possibility of designing an automated procedure for discharging a certain class of proof obligations that encode interesting security analysis problems for the class of access control schemas that we are defining.
3.4 The set $\Psi$ of state-change rules

A state-change rule is a formula of the form

$$\exists p, x, q. \text{knows}(p, x) \land \forall y, z, w. \text{msg}'(y, z, w) \iff \text{msg}(y, z, w) \lor (y = p \land z = \text{said}(x) \land w = q)$$  \hspace{1cm} (6)$$

that is usually abbreviated as (1). Intuitively, the unprimed and primed versions of $\text{msg}$ denote the state of the network immediately before and after, respectively, of the execution of the state-change rule. Let $S_1$ and $S_2$ be two states in $S$ and $\psi$ be a formula of the form (6), then $S_1 \rightarrow \psi S_2$ iff

$$S_2 := \{\text{msg}(y, z, w) \leftarrow (y = p \land z = \text{said}(x) \land w = q) \sigma \mid S_1 \cup \{\text{knows}(p, x) \sigma\} \text{ is satisfiable modulo } T_S \text{ for } \sigma \text{ ground substitution of } p, x, q\}.$$  

When $S_2 \neq \emptyset$, the state-change rule is enabled in $S_1$; otherwise (i.e., $S_2 = \emptyset$) it is disabled in $S_1$. This concludes the definition of our class of access control schema.

3.5 Reachability problems

In the class of ACSs defined above, policies rely on conditions that are determined by the exchange of messages (cf. predicate $\text{msg}$ and the CLP rule (3)) and non-mechanizable activities (cf. predicate $\text{uknows}$ and the CLP rule (2)). The state-change rules in $\Psi$ can only modify $\text{msg}$ and leave $\text{uknows}$ unconstrained since it is very difficult to model how humans decide to create a certain certificate. Returning to the CRO scenario, consider the assertion of fact (F3) as an example of a certificate that can be created at any time of the execution sequence of the system. To emphasize this aspect, we explicitly define the notion of (instance of) the reachability problem, although technically it can be derived from that of reachability problem given at the beginning of this section.

**Definition 3 (Instance of the reachability problem).** Given a set $Po$ of policies and a query $G$, an instance of the reachability problem amounts to establishing whether there exist an integer $n \geq 0$ and constraint (ground) facts $H_0(\text{uknows}_0), \ldots, H_{n-1}(\text{uknows}_{n-1})$ such that

$$\mathcal{R}_n \cup \{G(\text{knows}_n)\} \text{ is satisfiable modulo } T_S,$$

where $\mathcal{R}_0 := \text{lfp}\{\{H_0(\text{uknows}_0)\} \cup Po(\text{knows}_0)\}$, $\mathcal{R}_i \rightarrow \psi \mathcal{R}_{i+1}$, $\mathcal{R}_{i+1} := \text{lfp}\{H_{i+1}(\text{uknows}_{i+1})\} \cup Po(\text{knows}_{i+1})$, $\text{msg}_i$, $\text{uknows}_i$, and $\text{knows}_i$ denote uniquely renamed copies of $\text{msg}$, $\text{uknows}$, and $\text{knows}$, respectively, and $\alpha(s_i)$ is the formula obtained from $\alpha$ by replacing each occurrence of the symbol $s$ with the renamed copy $s_i$ (for $i = 0, \ldots, n+1$).

Intuitively, $\mathcal{R}_0$ is the initial knowledge of the principals computed from their internal knowledge and the (exhaustive) application of the policies without any
exchange of messages (recall that, initially, we assume that the network contains no messages). Then, $R_1$ is obtained from $R_0$ by first applying one of the available state-change rule ($R_{i+1}$) followed by the exhaustive application of the policies that allows each principal to possibly derive new knowledge from both the exchanged messages and their internal knowledge. The $R_i$’s for $i \geq 2$ can be similarly characterized.

When there exists a value of $n$ such that (7) holds, we say that $G$ is reachable; otherwise (i.e., when, for every $n \geq 0$, we have that (7) does not hold) we say that $G$ is unreachable. If a bound $\pi$ on $n$ is known, we talk of a bounded reachability problem (with bound $\pi$). Since the reachability problem is undecidable even without considering non-mechanizable facts (see [5] for details) in the rest of the paper, we prefer to focus on identifying restricted instances of the (bounded) reachability problem that are useful in practice and can be automatically solved.

4 Automated Analysis of Scenario-based Specifications

Web service technology supports the development of distributed systems using services built by independent parties. Consequently, service composition and coordination become an important part of the web service architecture. Usually, individual specifications of web services are complemented by scenario-based specifications so that not only the intentions of individual services but also their expected interaction sequences can be documented. Interestingly, as we will show below, scenario-based specifications can be exploited to automatically and efficiently analyze security properties despite the well-known fact that unforeseen interplays among individually secure services may open security holes in their composition. The idea is to associate a scenario with an instance of a bounded reachability problem and then consider only the sequences of state-change rules that are compatible with the scenario itself.

4.1 Scenarios and bounded reachability problem

In our framework, a scenario is composed of a finite set of principals, some sequences of state-change rules of finite length, and a query $G$ that encodes an availability or a security property. Since a state-change rule (1) to be enabled requires a principal to have some internal knowledge, this component of the scenario implicitly identifies a sequence $H_0(\text{uknows}_0),...,H_{n-1}(\text{uknows}_{n-1})$ of non-mechanizable facts where $n \geq 0$ is the length of the longest sequence of state-change rules.

Example 4. An example of informal specification of a scenario is the Message Sequence Chart (MSC) for the CRO depicted on the left of Fig. 2 where the $m_i$’s are the messages containing the utterances in the table on the right of the same figure. It is easy to find the instance of (1) that allows one to send each message $m_i$. The solid lines in the MSC impose an ordering on messages while dashed lines (called co-regions, see, e.g., [25]) do not. So, for example, CA can send the two
Fig. 2: A user-pull scenario for the CRO: Ed sends to CRep the certificates for a positive answer to the query $G := \text{knows(CRep, a2i(Ed, cans))}$

certificates (in messages $m_1$ and $m_2$) about the roles of Ed and Helen to Ed in any order and these two certificates as well as the one sent by Helen about granting the permission to store documents in CRep (in message $m_3$) can be received in any order by Ed. For the CRO, the query $G := \text{knows(CRep, a2i(Ed, cans))}$ encodes an availability property saying that the trusted user Ed can get the permission of storing the document in CRep. Since the length of sequence of state-change rules specified by the scenario is 6, we can build an instance of the bounded reachability problem with bound $n = 6$ and the following sequence of non-mechanizable facts: $H_0 := \text{knows(CA, a2i(Ed, ise))}$, $H_1 := \text{knows(CA, a2i(Helen, ish))}$, $H_2 := \text{knows(Helen, a2i(Ed, cans))}$, and $H_i := \text{true}$ for $i = 3, 4, 5$. Other sequences are compatible with the scenario above, we just picked one. Such sequences are finitely many and can be exhaustively enumerated.

4.2 Decidability of a class of instances of the reachability problem

It would be interesting to find conditions that guarantee the decidability of this kind of instances of the bounded reachability problem with a given sequence of non-mechanizable facts. Before doing this, we need to discuss the following four technical conditions on the substrate theory $T_5$.

First, the fact that there is a finite and known number of principals in any scenario can be formalized by requiring the substrate theory $T_5$ to be such that:

$$(\text{C1}) \quad T_5 \models \forall x. \bigvee_{c \in C} x = c \land \bigwedge_{c_1 \in C, c_2 \in C \setminus \{c_1\}} c_1 \neq c_2$$

where $C$ is a finite set of constants of sort $\text{Principal}$. This imposes that there are exactly $|C|$ principals.

The second condition concerns the form of the policies: (C2) for each CLP rule in $P_0$, all the variables in its body but not in its head range over the set $C$ of principals. For the fix-point computation required to solve an instance of the reachability problem, variables not occurring in the head of a CLP rule
but only in its body must be eliminated by a suitable (quantifier elimination) procedure (see, e.g., [17]). Assuming that such variables range only over the set $C$ of principals—cf. condition (C1)—it is possible to replace each one of them with the constants in $C$ and take disjunction.

The third and fourth conditions state respectively that: (C3) $T_3$ must be closed under sub-structures (see Section 7.1 in Appendix) and (C4) $T_5$ must be locally finite (see Section 7.1 in Appendix). Examples of effectively locally finite theories are the theory of an enumerated data-type or the theory of linear orders (cf., e.g., [24]) for more details). The last two (more technical) conditions allow us to reduce the satisfiability of a formula containing universal quantifiers (namely, those in the CLP rules) to the satisfiability of ground formulae by instantiating variables with finitely many (representative) terms. This implies the decidability of the satisfiability modulo $T_5$ of (7) (in the definition of reachability problem) provided that it is decidable to check the satisfiability modulo $T_5$ of ground formulae.

**Theorem 1.** Let $P_0$ be a finite set of policies, $G$ a query, and $H_0, \ldots, H_{n-1}$ a sequence of non-mechanizable facts ($n \geq 1$). If (C1), (C2), (C3), and (C4) are satisfied and the satisfiability modulo $T_5$ of ground formulae is decidable, then the instance of the bounded reachability analysis problem (with bound $n$, sequence $H_0, \ldots, H_{n-1}$ of non-mechanizable facts, and query $G$) is decidable.

The proof of this result uses previous work [24] (see Section 7.2 in the Appendix) and yields the correctness of the automated analysis technique in Fig. 3.

This is only a first step towards the design of a usable automated technique. In fact, at each iteration of the procedure, the solution of the bounded reachability problem (at step 1(c)) requires to compute a fix-point and to check the satisfiability modulo the substrate theory. Such activities can be computationally quite expensive and any means of reducing their number is obviously highly desirable.

### 5 A Reduction Technique

The main drawback of the procedure in Fig. 3 is step 1 that forces the enumeration of all sequences in $\Sigma$. Unfortunately, $\Sigma$ can be very large, e.g., there are 12 execution paths of CRO that are compatible with the MSC in Fig. 2. To overcome this problem, in the rest of this section, we design a reduction technique that allows for the parallel execution of a group of “independent” exchanges of messages so that several sequences of $\Sigma$ can be considered at the same time in one iteration of step 1 in the algorithm of Fig. 3. In this way, the number of fix-point computations and satisfiability checks may be significantly reduced. The key to this refinement is a compact representation of the set $\Sigma$ based on an adaptation of Lamport’s happened before relation $\rightsquigarrow$ [16]. There are many possible choices to describe $\Sigma$, ranging from MSCs (as done in the previous section) to BPEL workflows for web services augmented with access control information [8]. We have chosen $\rightsquigarrow$ as a starting point because it is at the same time simple and general, and simplifies the design of our reduction technique.
**Input**: a substrate theory $T$, a set $Po$ of policies, and a scenario $= (a$ finite set $C$ of principals, a set $\Sigma$ of sequences of state-change rules of finite length, a query $G$)

**Output**: $G$ is reachable/unreachable

**Assumptions**: (C1), (C2), (C3), and (C4) are satisfied

1. For each sequence $\sigma \in \Sigma$:
   (a) Determine the sequence $H_0, \ldots, H_{|\sigma|}$ of non-mechanizable facts that enables the corresponding sequence of instances of the state-change rule (1).
   (b) Build an instance of the bounded reachability problem with bound $|\sigma|$, the non-mechanizable facts of the previous step, and the given query $G$.
   (c) Try to solve the instance of the bounded reachability problem built at previous step.
   (d) If one of the instances at the previous step turns out to be solvable, return that the query $G$ is reachable.

2. Return that the query $G$ is unreachable (if step 1(d) is never executed).

Fig. 3: Automated Analysis of Scenario-based Specifications (interleaving semantics)

### 5.1 The Causality Relation

$\rightsquigarrow$ is a means of ordering a set $L$ of events based on the potential causal relationship of pairs of events in a concurrent system. Formally, $\rightsquigarrow$ is a partial order on $L$, i.e., it is irreflexive, ($l \not\rightsquigarrow l$ for $l \in L$), transitive (if $l_1 \rightsquigarrow l_2$ and $l_2 \rightsquigarrow l_3$ then $l_1 \rightsquigarrow l_3$ for $l_1, l_2, l_3 \in L$), and anti-symmetric (if $l_1 \rightsquigarrow l_2$ then $l_2 \not\rightsquigarrow l_1$ for $l_1, l_2 \in L$).

Two distinct events $l_1$ and $l_2$ are **concurrent** if $l_1 \not\rightsquigarrow l_2$ and $l_2 \not\rightsquigarrow l_1$ (i.e., they cannot causally affect each other). In the usual interleaving semantics, the set of possible executions can be seen as the set of all linear orders that extend $\rightsquigarrow$. Formally, $\rightsquigarrow_t$ is a linear extension of $\rightsquigarrow$ if $\rightsquigarrow_t$ is a total order (i.e., a partial order that is also total, for every $l_1, l_2 \in L$ we have that $l_1 \rightsquigarrow_t l_2$ or $l_2 \rightsquigarrow_t l_1$) that preserves $\rightsquigarrow$ (i.e., for every $l_1, l_2 \in L$, if $l_1 \rightsquigarrow l_2$, then $l_1 \rightsquigarrow_l l_2$).

For example, if $L = \{l_1, l_2\}$, $l_1$ and $l_2$ are concurrent, then both $l_1 \rightsquigarrow l_2$ and $l_2 \rightsquigarrow l_1$ are possible linear extensions since $l_1$ and $l_2$ cannot causally affects each other. Enumerating all the elements of the set $E(\rightsquigarrow_t)$ of linear extensions of the partial order $\rightsquigarrow$ can be done in $O(|E(\rightsquigarrow)|)$ constant amortized time [22] and computing $|E(\rightsquigarrow)|$ (i.e., counting the number of linear extensions of $\rightsquigarrow$) is #P-complete [9].

In our framework, $L$ is the set of instances of the state-change rule (1) considered in a scenario-based specification. Thus, the relation $\rightsquigarrow$ must be specialized so that the following two constraints must hold:

**(COMP1)** the enabledness of concurrent (according to $\rightsquigarrow$) instances of (1) must be preserved—i.e., two such instances should not to causally affect
each other by enabling (disabling) a disabled (enabled, respectively) state-change rule,

(COMP2) any execution of the concurrent events in a (finite) set \( L \), each of which causally affects another event \( l \) not in \( L \), results in a state in which \( l \) is enabled.

These requirements are formalized as follows.

(COMP1) if \( l_2 \) is enabled (disabled) in state \( S \) then it is still enabled (disabled, respectively) in state \( S' \) for \( S \rightarrow l_1 \rightarrow S' \) and the same must hold when swapping \( l_1 \) with \( l_2 \).

(COMP2) \( \text{Pre}(l') \subseteq L \) is such that \( l \rightsquigarrow^* l' \) for each \( l \in \text{Pre}(l') \) and there is no \( l'' \in L \setminus \text{Pre}(l') \) such that \( l'' \rightsquigarrow^* l' \), then \( l' \) is enabled in \( S' \) where \( S \rightarrow l_1 \rightarrow l_k \rightarrow S' \) and \( \text{Pre}(l') = \{l_1, ..., l_k\} \).

(COMP1) implies that the execution of either \( l_1 \) or \( l_2 \) followed by \( l_2 \) or \( l_1 \), respectively, will produce two identical states provided that the two executions start from the same initial state. (COMP2) says that once the action of sending a message is enabled, it persists to be so; this is related to the fact that of (1) can only add messages to \( \text{msg} \). Although (COMP2) seems to be restrictive at first sight, it is adequate for checking reachability (safety) properties as we do in this paper.

**Definition 4 (Causality Relation).** A partial order relation \( \rightsquigarrow \) on a finite set \( L \) of instances of (1) that satisfies (COMP1) and (COMP2) is called a causality relation.

The tuple \((C, L, \rightsquigarrow, G)\) identifies a scenario \((C, \Sigma, G)\) for \( C \) a finite set of principals, \( G \) a ground query, and \( \Sigma \) is the set of sequences obtained by enumerating all the linear extensions of \( \rightsquigarrow \) on \( L \). Since any linear extension of \( \rightsquigarrow \) is of finite length (as \( \rightsquigarrow \) is acyclic), we will also call \((C, L, \rightsquigarrow, G)\) a scenario.

We observe that when the state \( S \) is given, it is possible to show that both (COMP1) and (COMP2) are decidable (the proof is similar to that of Theorem 1). In practice, it is not difficult to argue that (COMP1) and (COMP2) hold for a given scenario.

**Example 5.** To illustrate this, we reconsider the scenario informally specified in Fig. 2 for the CRO and recast it in the formal framework developed above, as shown in Fig. 4.

There is an obvious correspondence between the entries of the tables in the two figures. The message \( m_1 \) is the result of executing \( SEC \), \( m_2 \) of \( SHC \), \( m_3 \) of \( SPC \), \( m_4 \) of \( SEC_2 \), \( m_5 \) of \( SHC_2 \), and \( m_6 \) of \( SPC_2 \). There is also a correspondence between the MSC in Fig. 2 and the causality relation \( \rightsquigarrow \) in Fig. 4.

Now, we show that the requirement (COMP1) holds for each pair \((l_1, l_2)\) of concurrent rule instances in \( L \) as follows. For \((SEC)\) and \((SHC)\), we have that if the latter is enabled (disabled) before the execution of the first \((SHC)\), it remains enabled (disabled, respectively) after its execution; the vice versa also
5.2 A reduction technique based on causality relations.

So far, we have shown that a causality relation can be exploited to compactly specify a scenario. Here, we show how it can be used to dramatically reduce the number of fix-point computation and satisfiability checks required by the analysis technique in Fig. 3 while preserving its completeness. The key idea is the following.

Since pairs of concurrent rule instances cannot causally affect each other, it is possible to execute them in parallel, i.e. adopting a partial order semantics. In fact, any linearization of the parallel execution, in the usual interleaving semantics, will yield the same final state obtained from the parallel execution. This has two advantages.

First, a single parallel execution of concurrent events correspond to a (possibly large) set of linear executions. Second, the length of the parallel execution is shorter than those of the associated linear executions. The number of fix-point computations and satisfiability checks needed to solve a bounded reachability problem can be reduced depending on the degree of independence of the rule instances in the scenario. The price to pay is a modification of the definition of

\[
C := \{\text{Ed, Helen, CA, CRep}\}
\]

\[
L := \{\text{SEC, SEC}_2, \text{SHC, SHC}_2, \text{SPC}\}
\]

\(\rightsquigarrow\) is the smallest partial order s.t.:

- \(\text{SEC} \rightsquigarrow \text{SEC}_2\)
- \(\text{SHC} \rightsquigarrow \text{SHC}_2\)
- \(\text{SPC} \rightsquigarrow \text{SPC}_2\)

| \(\text{(SEC)}\) & \(\text{knows(Ed, a2i(Ed, ise))} \Rightarrow \ominus\text{msg(Ed, said(a2i(Ed, ise)), Ed)}\) |
| \(\text{(SHC)}\) & \(\text{knows(Ed, a2i(Helen, ise))} \Rightarrow \ominus\text{msg(Ed, said(a2i(Helen, ise)), Ed)}\) |
| \(\text{(SPC)}\) & \(\text{knows(Helen, a2i(Ed, cans))} \Rightarrow \ominus\text{msg(Helen, said(a2i(Ed, cans)), Ed)}\) |
| \(\text{(SEC)}\) & \(\text{knows(Ed, s2i(CA, said(a2i(Ed, ise))))} \Rightarrow \ominus\text{msg(Ed, said(s2i(CA, said(a2i(Ed, ise))), CRep)}\) |
| \(\text{(SHC)}\) & \(\text{knows(Ed, s2i(CA, said(a2i(Helen, ise))))} \Rightarrow \ominus\text{msg(Ed, said(s2i(CA, said(a2i(Helen, ise))), CRep)}\) |
| \(\text{(SPC)}\) & \(\text{knows(Ed, s2i(Helen, said(a2i(Ed, cans))))} \Rightarrow \ominus\text{msg(Ed, said(s2i(Helen, said(a2i(Ed, cans))), CRep)}\) |

Fig. 4: Formalization of the CRO scenario in Fig. 2

holds. Similar observations hold also for the remaining pairs of concurrent events in \(L\).

Then, we show that the requirement \((\text{COMP2})\) holds for the events \((\text{SEC})\) and \((\text{SEC}_2)\) that are such that \((\text{SEC}) \rightsquigarrow (\text{SEC}_2)\). When \((\text{SEC})\) is executed, \((\text{SEC}_2)\) becomes enabled since the fact \(\text{msg(Ed, said(a2i(Ed, ise)), Ed)}\) holds as the result of executing \((\text{SEC})\): by the CLP rule (3) it is possible to derive \(\text{knows(Ed, s2i(CA, said(a2i(Ed, ise)))))}\) that is precisely the enabling condition of \((\text{SEC}_2)\). Similar observations hold for \((\text{SHC})\) and \((\text{SHC}_2)\) as well as \((\text{SPC})\) and \((\text{SPC}_2)\). Intuitively, \(\rightsquigarrow\) formalizes the obvious remark that, before Ed can forward a certificate to CRep (about his role, Helen’s role, or the permission to store documents), he must have preliminarily received it regardless of the order in which he has received the certificates from CA and Helen. \(\Box\)
reachability problem (cf. the end of Section 3) to adopt a partial order semantics.
We explain in more detail these ideas below.

Let $Po$ be a (finite) set of policies and $(C, L, \leadsto, G)$ a scenario, where $C$ is a
finite set of principals, $L$ is a finite set of rule instances of (1), $\leadsto$ is a causality
relation, and $G$ a query.

**Definition 5 (Reachability problem with partial-order semantics com-
patible with the causality relation $\leadsto$).** An instance of this problem amounts
to establishing whether there exist an integer $n \geq 0$ and (ground) constraint facts
$H_0(\text{knows}_0), \ldots, H_{n-1}(\text{knows}_{n-1})$ s.t. $R_n \cup \{G(\text{knows}_n)\}$ is satisfiable modulo $T_S$, where

\[
R_0 := \text{lfp}(\{H_0(\text{knows}_0)\} \cup \text{Po}(\text{knows}_0)), \quad R_{i+1} := \text{lfp}(R_i \cup \{H_{i+1}(\text{knows}_{i+1})\} \cup \text{Po}(\text{knows}_{i+1})), \quad \text{and } R_i \rightarrow_1 \cdots \rightarrow_k R_{i+1} \text{ for } l_1, \ldots, l_k \in L \text{ such that any pair } (l_a, l_b) \text{ is of concurrent events } (a, b = 1, \ldots, k \text{ and } a \neq b).
\]

Definition 5 is almost identical to to the Definition 3. The main difference
is in allowing the execution of a sequence $l_1, \ldots, l_k$ of exchange of messages pro-
vided that these are pairwise concurrent with respect to the causality relation.
Intuitively, we cumulate the effect of executing the instances $l_1, \ldots, l_k$ of (1) in a
single step so that each principal can derive more knowledge from the exchange
of several messages than the exchange of just one message as it was the case
with the definition of reachability problem in Section 3.

With this new definition of reachability problem, we propose a refinement in
Fig. 5 of the analysis technique in Fig. 3. The main differences between the two
techniques are the following. In input, the scenario is given by using the notion
of causality relation in order to exploit the new definition of reachability problem
with the partial order semantics. Then, instead of considering all the possible
linear extensions of $\leadsto$ (as in Fig. 3), sets of pairwise concurrent events for parallel
execution are computed by using the causality relation. The idea is to use the
Hasse diagram $CG(\leadsto)$, called the causality graph in the following, associated
with the causality relation $\leadsto$, i.e. the transitive reduction of the relation $\leadsto$ seen as an oriented graph.
The crucial observation is that concurrent events can be identified by looking at
those nodes that are not connected by a path in the causality graph. Formally,
we need the following notion. An element $l$ is minimal in $L$ with respect to $\leadsto$
iff there is no element $l' \in L$ such that $l' \leadsto l$. Since $L$ is finite, minimal elements
of $\leadsto$ must exist (this is a basic property of partial orders over finite sets).
In step 2, the rule instances labeling the minimal elements in $L$ with respect to
$\leadsto$, that correspond to nodes with no incoming edges in the causality graph,
are the only that require non-mechanizable facts for them to be enabled. In
fact, all the rule instances labeling non-minimal elements with respect to the
causality are enabled by the execution of one or more rule instances that label
ancestors nodes in the causality graph, because of (COMP2). This is why we
compute $H_0$ in step 3 while all the other sets of non-mechanizable facts are
vacuously set to $true$ in step 5. The rule instances in $L_{P_0}$ labeling the nodes in
$P_0$ are concurrent because of (COMP1). In step 4, we exploit this observation
to compute the other set of concurrent rule instances that can be executed in
parallel by modifying the causality graph: the nodes and the edges whose sources
**Input:** a substrate theory $T_S$, a set $Po$ of policies, and a scenario $(C,L,\rightarrow,G)$: a finite set $C$ of principals, a finite set $L$ of instances of (1), a causality relation $\rightarrow$, a query $G$

**Output:** $G$ is reachable/unreachable

**Assumptions:** $(C1)$, $(C2)$, $(C3)$, $(C4)$, $(COMP1)$, and $(COMP2)$ are satisfied

1. Let $CG(\rightarrow)$ be the causality graph associated to $\rightarrow$.
2. Compute the set $P_0$ of nodes in $CG(\rightarrow)$ with no incoming edges and $L_{P_0}$ be the set of rule instances of (1) in $L$ labeling the nodes in $P_0$.
3. Determine the set $H_0$ of non-mechanizable facts that enables all the rule instances in $L_{P_0}$.
4. Set $j = 0$ and while the set of nodes in $CG(\rightarrow)$ is non-empty do
   (a) Delete from $CG(\rightarrow)$ all the nodes in $P_j$ and the edges whose sources are in $P_j$ and increment $j$ by 1.
   (b) Compute the set $P_j$ of nodes in $CG(\rightarrow)$ with no incoming edges and $L_{P_j} \subseteq L$ be the set of rule instances labeling the nodes in $P_j$.
5. Build an instance of the bounded reachability problem with partial order semantics compatible with the causality relation $\rightarrow$ with bound $j$, sequence $H_0,H_1,...,H_j$ of non-mechanizable facts where $H_i := \text{true}$ for $i = 1,...,j$, and the input query $G$. At each step $i$ of the bounded reachability problem, the rule instances in $L_{P_j}$ must be used for parallel execution.
6. If the instance of the bounded reachability problem is solvable, then return that the query $G$ is reachable; otherwise, return that $G$ is unreachable.

Fig. 5: Automated Analysis of Scenario-based Specifications (partial order semantics)

are in $P_0$ are deleted from $CG(\rightarrow)$ so that a the set $P_i$ of nodes with no incoming edges can be identified. The rule instances in $L_{P_j}$ labeling the nodes in $P_i$ are the new concurrent events that can be executed in parallel and so on. The procedure eventually terminates when no more nodes are left in the causality graph. Then, in step 5, the new definition of bounded reachability problem compatible with the causality relation $\rightarrow$ can be exploited by using the sets $L_{P_0},...,L_{P_j}$ of rules instances to be executed in parallel. If the instance is solvable then the query $G$ is reachable, otherwise it is unreachable. The correctness of the refined analysis in Fig. 3 stems from the fact that, by definition of causality relation, there exists an execution in the interleaving semantics for the concurrent events executed in parallel—because of $(COMP1)$—and the execution of rule instances that must happen before (with respect to $\rightarrow$), enable the execution of those that happen afterwards—according to $(COMP2)$.

We briefly illustrate how the refined version of the automated analysis works on the scenario in Fig. 2 of the CRO. According to the causality relation in Fig. 4, $(SEC)$, $(SHC)$, and $(SPC)$ are minimal elements of $\rightarrow$ (step 2). Thus, the non-mechanizable fact $H_0$ enabling their execution is the conjunction of the following three facts: $\text{uknows(CA,a2i(Ed,ise))}$, $\text{uknows(CA,a2i(Helen,ish))}$, and $\text{uknows(Helen,a2i(Ed,cans))}$. Deleting the nodes labeled by $(SEC)$, $(SHC)$, and $(SPC)$ with the corresponding edges in the causality graph leaves us with a graph containing three isolated nodes labeled by $(SEC_2)$, $(SHC_2)$, and $(SPC_2)$.
that can be executed in parallel. As a consequence, the bound of the reachability problem is 2 in which, initially, the parallel execution of \((SEC, SHC),\) and \((SPC)\) is enabled because of the non-mechanizable facts in \(H_0\) while the parallel execution of \((SEC_2, SHC_2),\) and \((SPC_2)\) is enabled, in the following step, because of the three new certificates available in the net. Even in this simple example, the savings of the reduction technique are important: the two-step parallel execution corresponds to 6 interleavings executions that must be considered when using the technique in Fig. 3.

We have implemented a prototype of the procedure above in WSSMT [3] that uses the SMT solver Z3 [31] for fix-point computation and SMT solving. The time taken to analyze the scenario in Fig. 4 with this prototype is negligible; larger examples are discussed in [3].

6 Discussion

We presented an automated technique to analyze scenario-based specifications of access control policies in open and distributed systems that takes into account human activities. It uses an instance of CLP to express policies and trust relationships, and reduces the analysis problem to fix-point computations and satisfiability checks. The first contribution is the decidability of the analysis of scenario-based specifications of ACSs. The second contribution is a reduction technique that allows us to make the decidability result useful in practice.

There are three main lines of research that are related to our work. First, several logic-based frameworks (e.g., [19, 14, 7, 15, 1, 24]) have been proposed to specify and analyze authorization policies with conditions depending on the environment of the system in which they are enforced. In principle, it is possible to consider the conditions depending on the execution of human activities as part of the environment and then re-use the available specification and analysis techniques. The problem in doing this is that the conditions for the execution of human activities are not explicitly modeled in the system so that their applicability is unconstrained. This results in a dramatic increase of the search space that makes the application of the available technique difficult, if possible at all. We avoid this state-explosion problem by considering scenario-based specifications that allow one to focus on a small sub-set of the possible sequences of events, as explained in Section 5. It would be interesting to adapt the abduction techniques in [6, 15] to identify which non-mechanizable facts need to be generated for the executability of complex scenarios in which condition \((COMP2),\) about the “monotonicity” of the events (Section 5), does not hold.

The second line of research is related to workflow analysis in presence of authorization policies, e.g., [8, 28]. On the one hand, such works specify the workflow as a partial ordering on tasks that is similar to the causality relation introduced here. On the other hand, these works abstract away the data-flow so that there is no need to specify compatibility conditions on the causality relation (cf. \((COMP1)\) and \((COMP2)\) in Section 5) as we do here because of
the modelling of the exchange of messages among principals. Another difference is that the specification of authorization policies is reduced to a minimum in \cite{8, 28} so as to simplify the study of the completion problem, i.e., whether there exists at least one assignment of users to tasks that allow for the execution of the whole workflow. Instead, we focus on reachability problems and we model, besides authorization policies, also trust relationship among principals. It would be interesting to study the decidability of the completion problem also in our richer framework.

The third line of research concerns the development of (semi) formal techniques for the analysis of human interventions. In \cite{10, 27} the authors aim to determine how a task is executed by humans and what special factors are involved to accomplish the goal the task is supposed to achieve. This line of work is based on informal methods to identify and analyze human actions in contrast to our framework that is based on a logical formalism. In \cite{26} the authors use graphs and deterministic finite state automata to model and analyze human behaviors in critical systems. Although we share the proposed formal approach with them, our framework differs for the capability to analyze systems influenced by non predictable human activities, in contrast with those predefined for industrial material-handling processes. Interesting works in modeling and reasoning about human operators are, e.g., \cite{30, 13}, where the analysis is based on concurrent game structures, a formalism similar to the ACS we used in Section 3.

The accurate verification analysis and the decidability result we presented in this paper are the major difference that distinguishes our work from their.

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We assume some familiarity with the syntactic and semantic notions of first-order logic [12] and Constraint Datalog [17].

Let \( \Sigma \) be a signature, i.e., a collection of function and predicate symbols with their arities. A \( \Sigma(x) \)-expression is an expression (a term, an atom, a literal, or a formula) built out of the symbols in \( \Sigma \) where at most the variables in the sequence \( x \) may occur free. We write \( E(x) \) to emphasize that \( E \) is a \( \Sigma(x) \)-expression. For two disjoint sequences of variables \( x \) and \( y \), we write \( x, y \) to denote their concatenation.

A \( \Sigma \)-structure \( \mathcal{N} \) is a sub-structure of a \( \Sigma \)-structure \( \mathcal{M} \) iff the domain of \( \mathcal{N} \) is contained in the domain of \( \mathcal{M} \) and the interpretations of the symbols of \( \Sigma \) in \( \mathcal{N} \) are restrictions of the interpretations of these symbols in \( \mathcal{M} \). A class \( \mathcal{CL} \) of \( \Sigma \)-structures is closed under sub-structures iff for every structure \( \mathcal{M} \in \mathcal{CL} \), if \( \mathcal{N} \) is a substructure of \( \mathcal{M} \) then \( \mathcal{N} \in \mathcal{CL} \).

Let \( \Sigma \) and \( \Sigma' \) be two signatures such that \( \Sigma \subseteq \Sigma' \). If \( \mathcal{M} \) is a \( \Sigma' \)-structure, then \( \mathcal{M}|_{\Sigma} \) is the reduct of \( \mathcal{M} \) obtained from \( \mathcal{M} \) by forgetting the interpretations of the symbols in \( \Sigma' \setminus \Sigma \).

A \( \Sigma \)-theory \( T \) is a set of \( \Sigma \)-formulae, called axioms. A \( \Sigma \)-theory \( T \) identifies a class \( \text{Mod}(T) \) of \( \Sigma \)-structures that are models of all formulae in \( T \). For each theory \( T \) considered in the paper, we assume that \( \text{Mod}(T) \neq \emptyset \) and we then say that \( T \) is consistent.

A \( \Sigma \)-formula \( \varphi(x) \) is \( T \)-satisfiable iff there exists a \( \Sigma \)-structure \( \mathcal{M} \in \text{Mod}(T) \), also called a \( \Sigma \)-model, such that \( \mathcal{M} \models \exists x. \varphi(x) \). The satisfiability modulo theory \( T \) problem, in symbols, \( \text{SMT}(T) \), consists of establishing the \( T \)-satisfiability of any quantifier-free \( \Sigma \)-formula.

A \( \Sigma \)-theory \( T \) is locally finite if \( \Sigma \) is finite and, for every set of constants \( a \), there exist finitely many ground terms \( t_1, \ldots, t_k \), called representatives, such that for every ground \( (\Sigma \cup a) \)-term \( u \), we have \( \mathcal{T} \models u = t_i \) for some \( i \). If the representatives are effectively computable from \( a \) and \( t_i \) is computable from \( u \), then \( T \) is effectively locally finite. A theory \( T \) admits quantifier elimination if for an arbitrary formula \( \varphi(x) \), possibly containing quantifiers, one can compute a \( T \)-equivalent quantifier-free formula \( \varphi'(x) \).

A formula of the Bernays-Schönfinkel-Ramsey (BSR) class is of the form \( \exists x. \forall y. \varphi(x, y) \), where \( x, y \) are (disjoint) tuples of variables and \( \varphi \) is a quantifier-free formula built out of a signature containing only predicate and constant symbols (i.e., no function symbol occurs in \( \varphi \)). Formulae of the BSR class where \( x \) is empty are called universal, whereas when \( y \) is empty they are called existential. It is easy to show that any theory whose axioms are universal BSR formulae is effectively locally finite. Satisfiability of BSR formulae is well-known to be decidable [23].

Let \( T \) be a \( \Sigma \)-theory and \( \mathcal{R} \) a tuple of predicate symbols not in \( \Sigma \). A BSR(\( T \))-formula is a formula of the form \( \exists x. \forall y. \varphi(x, y) \), where \( \varphi \) is a quantifier-free
Theorem 2 ([24]). Let $T$ be an effectively locally finite $\Sigma$-theory whose class of models is closed under sub-structures, the SMT$(T)$ problem be decidable, and $R$ be a finite set of predicate symbols such that $\Sigma \cap R = \emptyset$. Then, the satisfiability of BSR$(T)$-formulae is decidable.

Let $T$ be a $\Sigma$-theory. A constraint Datalog rule is a formula of the form

$$\forall x, y. \xi(x, y) \land \bigwedge_{i=1}^{n} A_i(x, y) \rightarrow A_0(x),$$

also written as

$$A_0(x) \leftarrow A_1(x, y) \land \cdots \land A_n(x, y) \land \xi(x, y),$$

where $A_i$ is an atom for $i = 0, 1, \ldots, n$, $\xi(x, y)$ is a quantifier-free $\Sigma(x, y)$-formula, called the constraint of the rule, and $x, y$ are disjoint tuples of variables; when $n = 0$, the constraint Datalog rule is also called a constraint fact.

The non-ground Herbrand base of a set $LP$ of constraint Datalog rules is the set of constraint facts modulo equality. The non-ground immediate consequences operator $S_{LP}$ is defined over a collection of constraint facts $F$ as follows: $S_{LP}(F)$ contains all the constraint facts of the form $A_0(x) \leftarrow \xi(x, y)$ when $A_0(x) \leftarrow A_1(x, y) \land \cdots \land A_n(x, y) \land \xi(x, y)$ is in $LP$, $A_i \leftarrow \xi'_i$ in $F$ for $i = 1, \ldots, n$, and $\xi$ is logically equivalent (in $T$) to $\xi'_1 \land \cdots \land \xi'_n$, where it is implicitly assumed that the variables in the rule and those in the constraint facts have been renamed so as to make them pairwise disjoint. It is possible to show the existence of the least fix-point $lfp(LP)$ of $S_{LP}$, which may be infinite.

It is sometimes possible to show that $lfp(S_{LP})$ is finite. Let $T$ admit quantifier-elimination. Let $r_0(x) \leftarrow A_0(x, y) \land \xi_0(x, y)$ be a constraint Datalog rule and $r_i(x) \leftarrow \xi_i(x)$ be a constraint fact for $\xi_i$ a $\Sigma(x)$-quantifier-free formula, $k_i$ the arity of $r_i$, and $i = 1, \ldots, n$. A constraint rule application produces $m \geq 0$ facts of the form $r_0(x) \leftarrow \xi'_j(x)$ where $\xi'_j$ is a quantifier-free $\Sigma(x)$-formula for $j = 1, \ldots, m$ ($m \geq 0$) and $\bigvee_{j=1}^{m} \xi'_j$ is equivalent (by the elimination of quantifiers in $T$) to the formula

$$\exists y. \left( \bigwedge_{i=1}^{n} \xi_i(x_k) \land \xi_0(z) \right),$$

where $y$ is the tuple of variables occurring in the body of the rule but not in the head. The algorithm to compute the least fix-point of a set of constraint Datalog rules is given in Fig. 6. The function constrFP terminates when all derivable new facts are implied by previously derived facts so that the least fix-point is reached.

Theorem 3 ([24]). Let $T$ be an effectively locally finite theory that admits elimination of quantifiers. Then, constrFP terminates returning a finite set of constraint facts.
function constrFP(F, R)
1 results ← F; Changed ← true;
2 while Changed do
3 Changed ← false
4 foreach rule ∈ R do
5 foreach tuple of constraint facts constructed from results do
6 newares ← constraint facts obtained by constraint rule application between rule and tuple
7 foreach fact ∈ newares do
8 if (results ∉ fact) then results ← results ∪ {fact};
9 Changed ← true;
10 end
11 end
12 end
13 end
14 return results

Fig. 6: Least fix-point computation of constraint Datalog rules (adapted from [17]): F is a set of constraint facts and R is a set of constraint Datalog rules.

7.2 Proof of Theorem 1

Proof. By considering Definition 2 in Section 3, let ψ₁,...,ψₙ₋₁ be a sequence (possibly containing repetitions) of elements in Ψ. Since Ψ is finite, there are finitely many sequences of elements of length n. Thus, we can enumerate all of such sequences. Since the value of n is given together with the collection {H₀,...,Hₙ₋₁} of non-mechanizable facts, it is possible to compute the least fixed point of the set Rₙ of constraints in (7) by repeatedly invoking the function constrFP of Fig. 6 in Section 7.1 as follows. Let

R₀ := constrFP({I(msg₀), H₀(unknown₀)}, P₀(knows₀))

be the set of constraint facts, generated from the initial state of the system, where P₀ has been obtained from P₀ by replacing each constraint Datalog rule r with the set \{rσ\}_σ where σ ranges over the mappings that associate the variables occurring in the body of r but not in the head of r with the constants of sort Principal in C, satisfying the (C1) requirement.

By Theorem 3 in Section 7.1, the invocation to constrFP terminates returning the finite set R₀ of constraint facts. In fact, both I(msg₀) and H₀(unknown₀) are constraint facts, P₀(knows₀) is a set of constraint Datalog rules, and the substrate theory T₅ satisfy the (C4) requirement. Note that there is no need to eliminate quantifiers during a constraint rule application as all the variables occurring in the body of a constraint Datalog rule of P₀(knows₀) occur also in its head. We are thus entitled to conclude that R₀ and R₀ defined in (7) are
equal (modulo variable renaming). Then, let
\[ R_1 = \text{constrFP}(\{H_1(\text{uknows}_1)\}, \text{Eff}_1(R_0, \psi_1(\text{msg}_0, \text{msg}_1))) \]
\[ \cup \widehat{P_0}(\text{knows}_1)), \] where, for \( i = 0 \),
\[ \text{Eff}_i(F, \psi(\text{msg}_i, \text{msg}_{i+1})) = \{ \text{msg}_{i+1}(y, z, w) \leftarrow \text{Upd}_i(y, z, w)\sigma \mid \tilde{F} \cup \{G_i\sigma\} \text{ is } T_5\text{-satisfiable by considering } p_1, x_1, \ldots, p_m, x_m \text{as fresh constants}\}, \]
\[ \tilde{F} := \{ A_0 \leftrightarrow \xi \mid (A_0 \leftarrow \xi) \in F\}, \]
\[ \text{Upd}_i \text{ and } G_i \text{ are obtained from } \text{Upd} \text{ and } G \text{ by replacing } \text{msg} \text{ and } \text{uknows} \text{ with } \text{msg}_i \text{ and } \text{uknows}_i, \text{respectively.} \]

Note that \( \text{Eff}_i(F, \psi) \) is finite if \( F \) is so. Also, the satisfiability of formulae over a signature extended with fresh constants is decidable when the satisfiability problem of the theory over its original signature is decidable [29].\(^2\) Thus, since the \( T_5\)-satisfiability is decidable by assumption, it is also decidable to check whether \( \tilde{F} \cup \{G\sigma\} \) is \( T_5\)-satisfiable. Hence, we are entitled to conclude that \( \text{Eff}_i(R_0, \psi(\text{msg}_0, \text{msg}_1))) \) is finite since \( R_0 \) is so. It is now easy to see that \( \text{constrFP}(\{H_1(\text{uknows}_1)\}, \text{Eff}_i(R_0, \psi(\text{msg}_0, \text{msg}_1))) \cup \widehat{P_0}(\text{knows}_1)) \) terminates for reasons that are similar to those discussed for the computation of \( R_0 \). The only difference is in the constraint Datalog rules derived from \( \text{Eff}(R_0, \psi(\text{msg}_0, \text{msg}_1))) \) for which it is not difficult to verify that the variables occurring in the body also occur in the head; thereby making it unnecessary to eliminate quantifiers as for \( R_0 \). Thanks to basic properties of \( \text{constrFP} \) (see [17]), we can derive that \( R_1 \) is equal (modulo variable renaming) to \( R_1 \). By a straightforward induction, generalizing the previous observations on \( R_1 \), it is possible to show that \( R_i \) is a finite set of constraint facts equal (modulo variable renaming) to \( R_i \), for \( i \geq 2 \).

We are thus left with the problem of checking the \( T_5\)-satisfiability of the (finite) set \( R_n \) for constraint facts, for some \( n \geq 0 \). This can be done as follows. For each constraint fact \( A \leftarrow \xi \) in \( R_n \), form the formula \( \phi_A \) by taking the disjunction of all \( \xi_i \)'s for \( i \geq 0 \). Then, take the disjunction of all the formulae \( \phi_A \) built at the previous step and build the quantifier-free formula \( \varphi \). The satisfiability of the \( BSR(T_5)\)-formula \( \varphi \wedge E \) is indeed decidable thanks to Theorem 2 in Section 7.1. This concludes the proof.

### 7.3 Main derivation of CRO scenario

In this section, we illustrate the access control layer underlying (a simplified version of) an e-Government application, first described in [2]. We already have a description of the CRO case study in the body of the paper (see Section 1.2) and here we show, step by step, the derivation process of its main access control query \( \text{knows}(\text{CRep}, a2i(\text{Ed}, \text{cans})) \).

\(^2\)Note that we can extend the signatures with fresh (Skolem) constants since we consider all the classes of models of the substrate theory \( T_5 \). The least fix-point semantics is used only when considering the constraint Datalog rules specifying the policies, not the substrate theory.
A scenario of a possible run of the system is illustrated in Fig. 7(a) where, in the displayed state, the three certificates (C1), (C2) and (C3) are in Ed’s possession can be derived from the set of non-mechanizable facts as described in Example 2 of Section 2.2.\(^3\) This is the result of the sending of three messages \(m_1\), \(m_2\), and \(m_3\) (labeling the arrows in Fig. 7(a)) as a consequence of the threefold application of the state-change rule (1).

After the successful processing of a car registration request, Ed is willing to permanently store it in CRep. In order to do this, he should comply to the CRep policy that regulates access to its central database. The decision of CRep to grant or deny to employees of a CRO the right to store a processed request in the database is based on the rules (P1)–(P4) described in Example 3 of Section 2.3 that we report again, informally, as follows:

(P1) an employee of a CRO can store documents in the CRep, if the head of the CRO permits it.

Since the application of (P1) is based on valid certificates, further policy regulations must specify the trust relationships that allow CRep to validate the certificates in its possessions. Such rules are:

(P2) certificates signed by CA are trusted by anyone,

\(^3\) Some of the details of the figure will be made clear below, e.g., the numbering of the states and the fact that we only show the most interesting states of the system.
(P3) any certificate signed by CA and countersigned by any principal is trusted as one being signed by CA itself, and

(P4) concerning the certificates about the permission of storing documents, the head of a CRO is trusted by anyone.

In order to satisfy the access policies (P1)–(P4) of CRep, Ed is supposed to forward to CRep the certificates in its possession, after signing each one of them. We want to remark that we preserve the countersign action as described in the original scenario definition in [2]. As widely discussed in [4], the role of CRep is not purely passive but can potentially check the digital signatures of principals involved. Even if it is not relevant for our purposes in this paper, it is important to underline the capability of our formalization to model more complex classes of scenarios. After receiving Ed’s certificates, CRep should be able to grant him the right to store documents in its internal database using all the information in its possession.

We want to define the CRO scenario as an instance of the ACS in Definition 2. So, we proceed defining each element of the transition system.

Let $T_S$ the effectively locally finite substrate theory underlying the CRO as described in Section 3.1, we take $Const_p := \{Ed, Helen, CRep, CA\}$ and $Const_a := \{ise, ish, cans\}$ as the two (countably) sets of constants of sort $Principal$ and $Attribute$ to identify the four principals depicted in Fig. 7(a) and the attributes of being an employee, being an head and having the right to store a document in the database, needed to built the initial certificates. The elements of $Const_a$ depend on the application we are considering and to characterize this set as particular primitive elements (not created by the “function” $tdOn$) we have to add to the set $In$ the following axioms:

$$\forall x. \text{prim}(x) \rightarrow x = ise \lor x = ish \lor x = cans \quad \text{and}$$

$$ise \neq ish \land ise \neq cans \land ish \neq cans$$

as described in Section 3.1.

We proceed with our formalization considering the set $K$ of $\{\text{uknows}\}$-atoms and the set $M$ of $\{\text{msg}\}$-atoms to represent the state of the system according to Section 3.2. The initial situation is represented by the three non-mechanizable facts (as introduced in the Example 1):

\begin{align*}
(F1) & \quad \text{uknows}(CA, a2i(Ed, ise)) \\
(F2) & \quad \text{uknows}(CA, a2i(Helen, ish)) \\
(F3) & \quad \text{uknows}(Helen, a2i(Ed, cans)),
\end{align*}

generated by the (arbitrary) human activities of CA and Helen, in order to put into the system the credentials needed to fulfill the established goal. Considering the content of the network we define a state to be to be initial if $M = \{\emptyset\}$.

Let $Po$ be the set of $BSR(T_S)$-formulae (1), (2), (3), together with axioms described according to Section 3.1 and application-dependent CLP rules (P1)-(P4) of Example 3, where $p, q,$ and $r$ are variables of sort $Principal$ and $x$ is a
variable of sort Infon. Clauses (P2), (P3) and (P4), stating the trust relationship between the various principals, are required to derive the hypotheses of (P1) in combination with the use of rule (4), as we will see below.

We have now all the information to define the CRO case study as an instance of the ACS according to Definition 2.

Let $CRO = \langle \text{uknows}, \text{msg}, I, Po, \Psi \rangle$ be the CRO-ACS with substrate theory $T_S$, where:

- $\text{knows}(\text{CRep}, a_2i(\text{Ed}, \text{cans}))$ is the goal $G$ that must be satisfied,
- $(F1)$, $(F2)$, $(F3)$ are the non-mechanizable facts declared above,
- $I(\text{msg})$ represents the first-order formula describing the initial state,
- $Po$ is the set of $BSR(T_S)$-formulae according to the definition above,
- $\Psi$ is a (finite) set of state-change rules of the form of (6).

We can now easily describe the execution of the system representing and collecting the set $St$ of states, by means of the two predicate symbols $\text{uknows}$ and $\text{msg}$ which model the dynamic part of the access control, unlike the static one modeled by the first-order theory $T_S$ previously defined.

We will write $s := K | M$ for a generic state, to represent the set $K \cup M$.

The initial situation in Fig. 8 (that for convenience we report here again) can be formalized by the following initial (symbolic) state:

$$s_0 := \{(F1), (F2), (F3)\} \cup \emptyset.$$

Now, applying the rule (2) in Section 2.2 and computing the fixed point by the constrFP procedure described in Fig. 6, we derive the following constraint facts:

$$\text{knows}(\text{CRep}, a_2i(\text{Ed}, \text{cans})), \text{knows}(\text{CA}, a_2i(\text{Helen}, \text{ish})), \text{and knows}(\text{Helen}, a_2i(\text{Ed}, \text{cans})).$$

At this point, we are ready to obtain the state $s_3$ depicted in Fig. 7(a) by repeatedly applying (three times) the state-change rule (1). Considering the grounding substitution $\sigma_1 := \{p \mapsto \text{CA}, q \mapsto \text{Ed}, x \mapsto a_2i(\text{Ed}, \text{ise})\}$, we have the following instance of (1):

$$\text{knows}(\text{CA}, a_2i(\text{Ed}, \text{ise})) \implies \oplus \text{msg}(\text{CA}, \text{said}(a_2i(\text{Ed}, \text{ise})), \text{Ed}),$$

which is clearly enabled in $s_0$. The application’s effect of $\psi \sigma_1$ on the constrained facts just calculated in $s_0$ leads to

$$s_1 := K_0 \{\text{msg}(\text{CA}, \text{said}(a_2i(\text{Ed}, \text{ise})), \text{Ed})\},$$

26
where $K_0 = \{(F1), (F2), (F3)\}$.

It is not difficult to see that two further applications $\psi \sigma_2, \psi \sigma_3$ (where $\sigma_{1,2}$ are suitable ground substitutions) of (1) allow us to obtain state

$$s_3 := K_0 \left\{ \begin{array}{l}
\text{msg}(CA, \text{said}(a2i(Ed, ise)), Ed), \\
\text{msg}(CA, \text{said}(a2i(Helen, ish)), Ed), \\
\text{msg}(Helen, \text{said}(a2i(Ed, cans)), Ed),
\end{array} \right\},$$

which is the formal counterpart of the configuration depicted in Fig. 7.

It is also immediate to see that the repetitive application of the function constrFP in Fig. 6 to states $s_2$ and $s_3$ generates the following three facts (by repeatedly applying clause (3)):

knows(Ed, s2i(CA, said(a2i(Ed, ise)))),
knows(Ed, s2i(CA, said(a2i(Helen, ish)))),
knows(Ed, s2i(Helen, said(a2i(Ed, cans))))

representing the formal counterpart of certificates $(C1), (C2)$ and $(C3)$ in Fig. 7(b).

Applying again the reasoning introduced up to now, these last three facts can be used by $Ed$ to counter-sign the certificates and send them to $CRep$ as depicted in Fig. 9(a) (by appropriate instances $\psi \sigma_4, \psi \sigma_5, \psi \sigma_6$ of rule (1) with
\( \sigma_{4,5,6} \) suitable ground substitutions), thereby deriving the state

\[
s_6 := K_0 | M_3 \cup \\
\{ \text{msg}(Ed, \text{said}(s2i(CA, \text{said}(a2i(Ed, ise)))), CRep), \\
\text{msg}(Ed, \text{said}(s2i(CA, \text{said}(a2i(Helen, ish)))), CRep), \\
\text{msg}(Ed, \text{said}(s2i(Helen, \text{said}(a2i(Ed, cans)))), CRep) \}
\]

which is the formalization of the configuration in Fig. 9, where \( M_3 \) abbreviates the second component of the state \( s_3 \) above. At the end of each single application of the state-change rule (1), the function \( \text{constrFP} \) returns (by using clause (3)) the set of certificates \( (C4),(C5) \) and \( (C6) \) represented in Fig.9(b).

Once the application has reached the state \( s_6 \), it is possible for \( \text{CRep} \) to take the decision to grant or deny to \( Ed \) the permission to store the processed request in the database. To this end, we need to validate the certificates that are in possession of \( \text{CRep} \) against the chain of trust relationships represented by the Horn clauses \( (P2)\)–\( (P4) \). More specifically, first, we consider the trust relationships concerning the certificates about the roles of the principals \( (C1), (C2) \) and then the certificate about the permission to store documents in the database, \( (C3) \). Formally, this can be done by using the Horn clause (4) introduced in Section 2.3.

In the following, we describe which instances of (4) need to be considered and how their hypotheses are discharged. In order to positively answer the query \( G(\text{knows}(\text{CRep}, a2i(Ed, cans))) \), we consider the following instance of \( (P1) \):

\[
(G) \text{knows}(\text{CRep}, a2i(Ed, cans)) \leftarrow \text{knows}(\text{CRep}, a2i(Helen, ish)) \\
\wedge \text{knows}(\text{CRep}, a2i(Ed, ise)) \\
\wedge \text{knows}(\text{CRep}, s2i(Helen, \text{said}(a2i(Ed, cans)))).
\]

let us call it \( (G) \) (to recall the goal of the reachability analysis problem of Section 3). We have the problem of discharging the three hypotheses of \( (G) \), which can be grouped in two categories. In fact, the first two concern the roles of the principals (in particular, the fact that \( Ed \) should be an employee and that \( Helen \) be the head of the Car Registration Office), while the last is about the permission of storing documents in the central repository. We consider each category in detail.

**Validation of certificates about the roles of principals** Intuitively, we need to apply \( (P3) \) and \( (P2) \) so as to enable \( \text{CRep} \) to derive the pieces of knowledge that \( Ed \) is an employee (fact \( (H3) \) below) and that \( Helen \) is the head of the Car Registration Office (fact \( (H4) \) below). Indeed, in the derivation, the certificates \( (C4) \) and \( (C5) \) will be used which, in turn, are obtained from \( (C1) \) and \( (C2) \) via the applications of the state-change rule (1) described above. We begin by considering the following instances of (4):
\[(H1) \text{knows(CRep, } s2i(\text{CA}, \text{said(a2i(Ed, ise))))) } \leftarrow \\
\text{knows(CRep, } s2i(\text{Ed, said(s2i(\text{CA}, \text{said(a2i(Ed, ise))))})) \wedge \\
\text{knows(CRep, } a2i(\text{Ed, tdOn(s2i(\text{CA}, \text{said(a2i(Ed, ise))))})))
\]

\[(H2) \text{knows(CRep, } s2i(\text{CA}, \text{said(a2i(Helen, ish))))) } \leftarrow \\
\text{knows(CRep, } s2i(\text{Ed, said(s2i(\text{CA}, \text{said(a2i(Helen, ish)))))))) \wedge \\
\text{knows(CRep, a2i(\text{Ed, tdOn(s2i(\text{CA}, \text{said(a2i(Helen, ish))))}))}
\]

and notice that the first hypotheses are identical to \((C4)\) and \((C5)\) respectively. The second hypotheses of the two instances above are identical to the following two instances of \((P3)\) \text{knows(CRep, } a2i(\text{Ed, tdOn(s2i(\text{CA}, \text{said(a2i(Ed, ise))))})) \text{ and knows(CRep, } a2i(\text{Ed, tdOn(s2i(\text{CA}, \text{said(a2i(Helen, ish))))}))). So, we are entitled to consider the heads of the two instances of \((4)\) above as derived ground facts; let us call them \((H1)\) and \((H2)\), respectively. Then, consider two more instances of \((4)\):

\[(H3) \text{knows(CRep, } a2i(\text{Ed, ise)) } \leftarrow \\
\text{knows(CRep, } s2i(\text{CA}, \text{said(a2i(Ed, ise)))) \wedge \\
\text{knows(CRep, } a2i(\text{CA, tdOn(a2i(Ed, ise)))))
\]

\[(H4) \text{knows(CRep, } a2i(\text{Helen, ish)) } \leftarrow \\
\text{knows(CRep, } s2i(\text{CA, said(a2i(Helen, ish)))) \wedge \\
\text{knows(CRep, } a2i(\text{CA, tdOn(a2i(Helen, ish))))
\]

The first hypotheses of these two ground Horn clauses are identical to \((H1)\) and \((H2)\) respectively.

Their second hypotheses are identical to the following two instances of \((P2)\) \text{knows(CRep, } a2i(\text{CA, tdOn(a2i(Ed, ise)))) and knows(CRep, } a2i(\text{CA, tdOn(a2i(Helen, ish))))). As a consequence, we can consider the heads of the last two instances of \((4)\) as derived ground facts; let us call them \((H3)\) and \((H4)\), respectively.

Validation of certificates about the permission of storing documents Intuitively, we need to apply \((P4)\) so as to make immediately available to CRep the knowledge about the authorization state concerning the fact that Ed is permitted to store documents in the central repository by Helen (fact \((H6)\) below). We begin by considering the following instance of \((P4)\):

\[(H5) \text{knows(CRep, } a2i(\text{Ed, tdOn(s2i(Helen, said(a2i(Ed, cans))))))} \\
\text{ } \leftarrow \text{knows(CRep, } a2i(\text{Helen, ish}))
\]
Since the hypothesis of this Horn clause is identical to \((H4)\), we are entitled to consider the head of this clause as a derived fact; let us call it \((H5)\). Then, consider the following instance of (4):

\[
(H6) \text{knows}(\text{CRep}, s2i(\text{Helen}, \text{said}(a2i(\text{Ed}, \text{cans})))) \leftarrow \\
\text{knows}(\text{CRep}, s2i(\text{Ed}, \text{said}(s2i(\text{Helen}, \text{said}(a2i(\text{Ed}, \text{cans})))))))) \\
\wedge \\
\text{knows}(\text{CRep}, a2i(\text{Ed}.\text{tdOn}(s2i(\text{Helen}, \text{said}(a2i(\text{Ed}, \text{cans}))))).)
\]

The first hypothesis of the instance is identical to \((C6)\) and the second hypothesis is equal to \((H5)\); thus, we can consider its head as a derived ground fact, let us call it \((H6)\).

**Putting things together** At this point, it is sufficient to observe that the first hypothesis of \((G)\) is \((H4)\), the second is \((H3)\), and the last is \((H6)\) so that we can consider the head of the rule as a derived ground fact which is precisely the query that we were interested to answer, \text{knows}(\text{CRep}, a2i(\text{Ed}, \text{cans}))\).

As a final remark, we observe that it is possible to model several scenarios by considering alternative ways of distributing the certificates among the principals. For example, initially, only the certificate about his role can be sent to \text{Ed}, that concerning the role of \text{Helen} is sent to her, and that about the permission for \text{Ed} to store documents in the central repository can be sent to \text{CRep} directly from \text{Helen}. Indeed, this changes the way in which we can derive the query of interest as certificates are counter-signed by different principals so that trust relationships must be chained differently.

### 7.4 Dkal implementation for CRO scenario

In this section, we give a concrete implementation of the proposed CRO scenario in Section 1.2, by using the DKAL distributed authorization policy language provided by Microsoft Research [21]. The DKAL project page [11] contains a downloadable engine (implemented in F#) for running and checking DKAL policies. We implemented and successfully tested the scenario proposed in this paper, and we give the corresponding code in the following:

| Input specification code |
|--------------------------|
| type Principal = Dkal.Principal  |
| type Infon = Dkal.Infon         |
| type Attribute = System.String |
| type Evidence = Dkal.Evidence   |
| relation hasrole(P: Principal, A: Attribute) |
| relation haspermission(P: Principal, A: Attribute) |
---crep------------
//crep's policy
//rule to learn every justified infon comes from the communication
with X: Infon, E: Evidence
upon X [E]
do learn X [E]

//(P1) rule
with P: Principal, Q: Principal, R: Principal,
E: Evidence, E1: Evidence, E2: Evidence
if hasrole(P, ish) [E]
    if hasrole(Q, ise) [E1]
        if R said haspermission(Q, cans) [E2]
do once send to ed: haspermission(P, "ok ed, u can write!")

//(P2) rule
with P: Principal, X: Infon, E: Evidence
upon theca said X [E]
do learn X

//(P3) rule
with P: Principal, X: Infon, E: Evidence
upon P said theca said X [E]
do learn theca said X

//(P4) rule
with P: Principal, Q: Principal, R: Principal, E: Evidence, E1: Evidence
if hasrole(P, ish) [E]
learn hasrole(P, ish) -> Q said P said hasrole(R, ise) [E1]

---ed------------
//ed's policy
with P: Principal, Q: Principal, R: Principal, A: Attribute, E: Evidence, X: Infon
upon X -> R said haspermission(P, A) [E]
do once send to crep: Me said haspermission(Me, A) [E]

with Q: Principal, R: Principal, A: Attribute, E: Evidence, X: Infon
upon Q said hasrole(R, A) [E]
do once say with justification to crep: Q said hasrole(R, A) [E]

---theca----------
//theca's policy
// internal knowledge about roles of principals

substrate xml("<roleAssignments>
<roleAssignment id='ed' role='ise' />
<roleAssignment id='helen' role='ish' />
</roleAssignments>")
namespaces "roleAssignments"

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with P: Principal, P1: Principal, A: Attribute, A1: Attribute
if asInfon({| "roleAssignments" | ”//roleAssignment[@role='ise']/@id" | P |}) &&
asInfon({| "roleAssignments" | ”//roleAssignment[@id='ed']/@role" | A |}) &&
asInfon({| "roleAssignments" | ”//roleAssignment[@role='ish']/@id" | P1 |}) &&
asInfon({| "roleAssignments" | ”//roleAssignment[@id='helen']/@role" | A1 |})
do say with justification to P: hasrole(P, A)
say with justification to P: hasrole(P1, A1)

---helen----------------
// helen's policy
// internal knowledge about permissions of principals

substrate xml("<permissionAssignments>
<permissionAssignment id='ed' perm='cans' />
</permissionAssignments>")
namespaces "permissionAssignments"

with P: Principal, A: Attribute
if asInfon({| "permissionAssignments" | ”//permissionAssignment[@perm='cans']/@id" | P |}) &&
asInfon({| "permissionAssignments" | ”//permissionAssignment[@id='ed']/@perm" | A |})
do send with justification to P:
// delegation to Ed to say haspermission condition must be entailed by Ed's
// infostrate

with Q: Principal, A1: Attribute
Q said haspermission(Q, A1) -> Me said haspermission(P, A)

Output specification code

>> From theca to ed:
theca said hasrole(ed, "ise") [ signed by theca 1693769001 ] &&
theca said hasrole(helen, "ish") [ signed by theca 1939556616 ]

>> From helen to ed:
with Q: Dkal.Principal, A1: System.String
  Q said haspermission(Q, A1) ->
    helen said haspermission(ed, "cans")
    [ signed by helen -973755704 ]

>> From ed to crep:
ed said haspermission(ed, "cans") [ signed by helen -973755704 ] &&
ed said theca said hasrole(ed, "ise") [ signed by theca 1693769001 ]
  [ signed by ed -855343925 ] &&
ed said theca said hasrole(helen, "ish") [ signed by theca 1939556616 ]
  [ signed by ed -14984535 ]

>> From crep to ed
ok ed, u can write!

Fixed-point reached