On Secure Distributed Implementations of Dynamic Access Control

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Abstract
Distributed implementations of access control abound in distributed storage protocols. While such implementations are often accompanied by informal justifications of their correctness, our formal analysis reveals that their correctness can be tricky. In particular, we discover several subtleties in a state-of-the-art implementation based on capabilities, that can undermine correctness under a simple specification of access control.

We consider both safety and security for correctness; loosely, safety requires that an implementation does not introduce unspecified behaviors, and security requires that an implementation preserves the specified behavioral equivalences. We show that a secure implementation of a static access policy already requires some care in order to prevent unspecified leaks of information about the access policy. A dynamic access policy causes further problems. For instance, if accesses can be dynamically granted then the implementation does not remain secure—it leaks information about the access policy. If accesses can be dynamically revoked then the implementation does not even remain safe. We show that a safe implementation is possible if a clock is introduced in the implementation. A secure implementation is possible if the specification is accordingly generalized.

Our analysis shows how a distributed implementation can be systematically designed from a specification, guided by precise formal goals. While our results are based on formal criteria, we show how violations of each of those criteria can lead to real attacks. We distill the key ideas behind those attacks and propose corrections in terms of useful design principles. We show that other stateful computations can be distributed just as well using those principles.

1 Introduction
In most file systems, protection relies on access control. Usually the access checks are local—the file system maintains an access policy that specifies which principals may access which files, and any access to a file is guarded by a local check that enforces the policy for that file. In recent file systems, however, the access checks are distributed, and access control is implemented via cryptographic techniques. In this paper, we try to understand the extent to which these distributed implementations of access control preserve the simple character of local access checks.
We focus on implementations that appear in file systems based on networked storage [13]. In such systems, access control and storage are parallelized to improve performance. Execution requests are served by storage servers; such requests are guided by access requests that are served elsewhere by access-control servers. When a user requests access to a file, an access-control server certifies the access decision for that file by providing the user with an unforgeable capability. Any subsequent execution request carries that capability as proof of access; a storage server can efficiently verify that the capability is authentic and serve the execution request.

We formally study the correctness of these implementations vis-à-vis a simple specification of local access control. Implementing static access policies already requires some care in this setting; dynamic access policies cause further problems that require considerable analysis. We study these cases separately in Sections 2 and 3. Based on our analysis, we develop formal models and proofs for an implementation of arbitrary access policies in Section 6.

We consider both safety and security for correctness; loosely, safety requires that an implementation does not introduce unspecified behaviors, and security requires that an implementation preserves the specified behavioral equivalences. Our proofs of safety and security are built modularly by showing simulations; we develop the necessary definitions and proof techniques in Section 4.

Our analysis shows how a distributed implementation can be systematically designed from a specification, guided by precise formal goals. We justify those goals by showing how their violations can lead to real attacks (Sections 2 and 3). Further, we distill the key ideas behind those attacks and propose corrections in terms of useful design principles. We show that other stateful computations can be distributed just as well using those principles (Section 7).

Comparison with related work This paper culminates a line of work that we begin in [10] and continue in [11]. In [10], we show how to securely implement static access policies with capabilities; in [11], we present a safe (but not secure) implementation of dynamic access policies in that setting. In this paper, we carefully review those results, and systematically analyze the difficulties that arise for security in the case of dynamic access policies. Our analysis leads us to develop variants of the implementation in [11] that we can prove secure with appropriate assumptions. The proofs are built by a new, instructive technique, which may be of independent interest.

Further, guided by our analysis of access control, we show how to automatically derive secure distributed implementations of other stateful computations. This approach is reminiscent of secure program partitioning [22].

Access control for networked storage has been studied in lesser detail by Gobioff [13] using belief logics, and by Halevi et al. [15] using universal composability [9]. The techniques used in this paper are similar to those used by Abadi et al. for secure implementation of channel abstractions [2] and authentication primitives [3], and by Maffeis to study the equivalence of communication patterns in distributed query systems [17]. These techniques rely on programming languages concepts, including testing equivalence [21] and full abstraction [19][11]. A huge body of such techniques have been developed for formal specification and verification of systems.
We do not consider access control for untrusted storage [16] in this paper. In file systems based on untrusted storage, files are cryptographically secured before storage, and their access keys are managed and shared by users. As such, untrusted storage is quite similar to public communication, and standard techniques for secure communication on public networks apply for secure storage in this setting. Related work in that area includes formal analysis of protocols for secure file sharing on untrusted storage [18, 8], as well as correctness proofs for the cryptographic techniques involved in such protocols [7, 12, 6].

2 Review: the case of static access policies

To warm up, let us focus on implementing access policies that are static. In this case, a secure implementation already appears in [10]. Below we systematically reconstruct that implementation, focusing on a detailed analysis of its correctness. This analysis allows us to distill some basic design principles, marked with bold $R$, in preparation for later sections, where we consider the more difficult problem of implementing dynamic access policies.

Consider the following protocol, $NS^*$, for networked storage. Principals include users $U, V, W, \ldots$, an access-control server $A$, and a storage server $S$. We assume that $A$ maintains a (static) access policy $F$ and $S$ maintains a store $\rho$. Access decisions under $F$ follow the relation $F \vdash_U op$ over users $U$ and operations $op$. Execution of an operation $op$ under $\rho$ follows the relation $\rho[\llbracket op \rrbracket] \Downarrow \rho'[\llbracket r \rrbracket]$ over next stores $\rho'$ and results $r$. Let $K_{AS}$ be a secret key shared by $A$ and $S$, and $\text{mac}$ be a function over messages and keys that produces unforgeable message authentication codes (MACs) [14]. We assume that MACs can be decoded to retrieve their messages. (Usually MACs are explicitly paired with their messages, so that the decoding is trivial.)

\begin{align*}
(1) & \quad U \rightarrow A : op \\
(2) & \quad A \rightarrow U : \text{mac}(op, K_{AS}) \text{ if } F \vdash_U op \\
(2') & \quad A \rightarrow U : \text{error} \text{ otherwise} \\
(3) & \quad V \rightarrow S : \kappa \\
(4) & \quad S \rightarrow V : r \text{ if } \kappa = \text{mac}(op, K_{AS}) \\
& \quad \quad \quad \quad \text{ and } \rho[\llbracket op \rrbracket] \Downarrow \rho'[\llbracket r \rrbracket] \\
(4') & \quad S \rightarrow V : \text{error} \text{ otherwise}
\end{align*}

Here a user $U$ requests $A$ for access to an operation $op$, and $A$ returns a capability for $op$ only if $F$ specifies that $U$ may access $op$. Elsewhere, a user $V$ requests $S$ to execute an operation by sending a capability $\kappa$, and $S$ executes the operation only if $\kappa$ authorizes access to that operation.

What does “safety” or “security” mean in this setting? A reasonable specification of correctness is the following trivial protocol, $IS^*$, for ideal storage. Here principals include users $U, V, W, \ldots$ and a server $D$. The access policy $F$ and the store $\rho$ are

\footnote{By convention, we use superscripts $^*$ and $d$ to denote “static” and “dynamic”, and superscripts $+$ and $-$ to denote “extension” and “restriction.”}
both maintained by \(D\); the access and execution relations remain as above. There is no cryptography.

\[
\begin{align*}
(i) & \quad V \rightarrow D : op \\
(ii) & \quad D \rightarrow V : r \text{ if } F \vdash_V op \text{ and } \rho[[op]] \Downarrow \rho'[r] \\
(i'i) & \quad D \rightarrow V : \text{error otherwise}
\end{align*}
\]

Here a user \(V\) requests \(D\) to execute an operation \(op\), and \(V\) executes \(op\) only if \(F\) specifies that \(V\) may access \(op\). This trivial protocol is correct “by definition”; so if \(NS^s\) implements this protocol, it is correct as well.

What notions of implementation correctness are appropriate here? A basic criterion is that of safety \([4]\).

**Definition 1 (Safety).** Under any context (adversary), the behaviors of a safe implementation are included in the behaviors of the specification.

In practice, a suitable notion of inclusion may need to be crafted to accommodate specific implementation behaviors by design (such as those due to messages (1), (2), and (2’) in \(NS^s\)). Typically, those behaviors can be eliminated by a specific context (called a “wrapper”), and safety may be defined modulo that context as long as other, interesting behaviors are not eliminated.

Still, safety only implies the preservation of certain trace properties. A more powerful criterion may be derived from the programming languages concept of semantics preservation, otherwise known as full abstraction \([19, 1]\).

**Definition 2 (Security).** A secure implementation preserves behavioral equivalences of the specification.

In this paper, we tie security to an appropriate may testing congruence \([21]\). We consider a protocol instance to include the file system and some code run by “honest” users, and assume that an arbitrary context colludes with the remaining “dishonest” users. From any \(NS^s\) instance, we derive its \(IS^s\) instance by an appropriate refinement map \([3]\). If \(NS^s\) securely implements \(IS^s\), then for all \(NS^s\) instances \(Q_1\) and \(Q_2\), \(Q_1\) and \(Q_2\) are congruent if their \(IS^s\) instances are congruent.

Security implies safety for all practical purposes, so a safety counterexample usually suffices to break security. For instance, we are in trouble if operations that cannot be executed in \(IS^s\) can somehow be executed in \(NS^s\) by manipulating capabilities. Suppose that \(F \not\vdash_V op\) for all dishonest \(V\). Then no such \(V\) can execute \(op\) in \(IS^s\). Now suppose that some such \(V\) requests execution of \(op\) in \(NS^s\). We know that \(op\) is executed only if \(V\) shows a capability \(\kappa\) for \(op\). Since \(\kappa\) cannot be forged, it must be obtained from \(A\) by some honest \(U\) that satisfies \(F \vdash_U op\). Therefore:

**R1** Capabilities obtained by honest users must not be shared with dishonest users.

(However \(U\) can still share \(\kappa\) with honest users, and any execution request with \(\kappa\) can then be reproduced in the specification as an execution request by \(U\).)

While (R1) prevents explicit leaking of capabilities, we in fact require that capabilities do not leak any information that is not available to \(IS^s\) contexts. Information may also be leaked implicitly (by observable effects). Therefore:
R2 Capabilities obtained by honest users must not be examined or compared.

Both (R1) and (R2) may be enforced by typechecking the code run by honest users.

Finally, we require that information is not leaked via capabilities obtained by dishonest users. (Recall that such capabilities are already available to the adversary.) Unfortunately, a capability for an operation \( op \) is provided only to those users who have access to \( op \) under \( F \); in other words, \( A \) leaks information on \( F \) whenever it returns a capability! This leak breaks security. Why? Consider implementation instances \( Q_1 \) and \( Q_2 \) with \( op \) as the only operation, whose execution returns \textit{error} and may be observed only by honest users; suppose that a dishonest user has access to \( op \) in \( Q_1 \) but not in \( Q_2 \). Then \( Q_1 \) and \( Q_2 \) can be distinguished by a context that requests a capability for \( op \)—a capability will be returned in \( Q_1 \) but not in \( Q_2 \)—but their specification instances cannot be distinguished by any context.

Why does this leak concern us? After all, we expect that executing an operation \textit{should} eventually leak some information about access to that operation, since otherwise, having access to that operation is useless! However the leak here is premature; it allows a dishonest user to obtain information about its access to \( op \) in an undetectable way, \textit{without} having to request execution of \( op \).

To prevent this leak, we must modify the protocol:

R3 “Fake” capabilities for \( op \) must be returned to users who do not have access to \( op \).

The point is that it should not be possible to distinguish the fake capabilities from the real ones prematurely. Let \( K_{AS} \) be another secret key shared by \( A \) and \( S \). As a preliminary fix, let us modify the following message in \( NS^s \).

\[
\text{(2')} A \rightarrow U : \text{mac}(op, K_{AS}) \quad \text{if } F \uparrow_U op
\]

Unfortunately this modification is not enough, since the adversary can still compare capabilities that are obtained by different users for a particular operation \( op \), to know if their accesses to \( op \) are the same under \( F \). To prevent this leak:

R4 Capabilities for different users must be different.

For instance, a capability can mention the user whose access it authenticates. Making the meaning of a message explicit in its content is a good design principle for security \cite{5}, and we use it on several occasions in this paper. Accordingly we modify the following messages in \( NS^s \).

\[
\text{(2)} A \rightarrow U : \text{mac}((U, op), K_{AS}) \quad \text{if } F \vdash_U op \\
\text{(2')} A \rightarrow U : \text{mac}((U, op), K_{AS}) \quad \text{otherwise}
\]

\[
\text{(4)} S \rightarrow V : r \quad \text{if } \kappa = \text{mac}((\None, op), K_{AS}) \quad \text{and } \rho[\text{op}] \Downarrow \rho'[r]
\]

(On receiving a capability \( \kappa \) from \( V \), \( S \) still does not care whether \( V \) is the user to which \( \kappa \) is issued, even if that information can now be obtained from \( \kappa \).)

The following result can then be proved (cf. \cite{10}).

\textbf{Theorem 1.} \( NS^s \) securely implements \( IS^s \).
3 The case of dynamic access policies

We now consider the more difficult problem of implementing dynamic access policies. Let \( F \) be dynamic; the following protocol, \( NS^d \), is obtained by adding administration messages to \( NS^* \). Execution of an administrative operation \( \theta \) under \( F \) follows the relation \( F[\theta] \Downarrow F'[\theta] \) over next policies \( F' \) and results \( r \).

\[
\begin{align*}
(5) & \quad W \rightarrow A : \theta \\
(6) & \quad A \rightarrow W : r \text{ if } F \vdash W \theta \text{ and } F[\theta] \Downarrow F'[\theta] \\
(6') & \quad A \rightarrow W : \text{error otherwise}
\end{align*}
\]

Here \( A \) executes \( \theta \) (perhaps modifying \( F \)) if \( F \) specifies that \( W \) controls \( \theta \). The following protocol, \( IS^d \), is obtained by adding similar messages to \( IS^* \).

\[
\begin{align*}
(iii) & \quad W \rightarrow D : \theta \\
(iv) & \quad D \rightarrow W : r \text{ if } F \vdash W \theta \text{ and } F[\theta] \Downarrow F'[\theta] \\
(iv') & \quad D \rightarrow W : \text{error otherwise}
\end{align*}
\]

Unfortunately \( NS^d \) does not remain secure with respect to \( IS^d \). Consider the \( NS^d \) pseudo-code below. Informally, acquire \( \kappa \) means “obtain a capability \( \kappa \)” and use \( \kappa \) means “request execution with \( \kappa \)” ; chmod \( \theta \) means “request access modification \( \theta \)” ; and success means “detect successful use of a capability”. Here \( \kappa \) is a capability for an operation \( op \) and \( \theta \) modifies access to \( op \).

\[
\begin{align*}
t_1 & \quad \text{acquire} \kappa; \text{chmod} \theta; \text{use} \kappa; \text{success} \\
t_2 & \quad \text{chmod} \theta; \text{acquire} \kappa; \text{use} \kappa; \text{success}
\end{align*}
\]

Now \( t_1 \) and \( t_2 \) map to the same \( IS^d \) pseudo-code chmod \( \theta \); exec \( op \); success—informally, exec \( op \) means “request execution of \( op \)” . Indeed, requesting execution with \( \kappa \) in \( NS^d \) amounts to requesting execution of \( op \) in \( IS^d \), so the refinement map must erase instances of acquire and replace instances of use with the appropriate instances of exec. However, suppose that initially no user has access to \( op \), and \( \theta \) specifies that all users may access \( op \). Then \( t_1 \) and \( t_2 \) can be distinguished by testing the event success. In \( t_1 \) \( \kappa \) does not authorize access to \( op \), so success must be false; but in \( t_2 \) \( \kappa \) may authorize access to \( op \), so success may be true.

Moreover, if revocation is possible, \( NS^d \) does not even remain safe with respect to \( IS^d \)! Why? Let \( \theta \) specify that access to \( op \) is revoked for some user \( U \), and revoked be the event that \( \theta \) is executed (thus modifying the access policy). In \( IS^d \), \( U \) cannot execute \( op \) after revoked. But in \( NS^d \), \( U \) can execute \( op \) after revoked by using a capability that it acquires before revoked.

Safety in a special case One way of eliminating the counterexample above is to make the following assumption:

A1 Accesses cannot be dynamically revoked.

We can then prove the following new result (see Section 6).
Theorem 2. \( NS^d \) safely implements \( IS^d \) assuming (A\( ^1 \)).

The key observation is that with (A\( ^1 \)), a user \( U \) cannot access \( op \) until it can always access \( op \), so \( U \) gains no advantage by acquiring capabilities early.

Safety in the general case  Safety breaks with revocation. However, we can recover safety by introducing time. Let \( A \) and \( S \) share a logical clock (or counter) that measures time, and let the same clock appear in \( D \). We have that:

**R5** Any capability that is produced at time \( \text{Clk} \) expires at time \( \text{Clk} + 1 \).

**R6** Any administrative operation requested at time \( \text{Clk} \) is executed at the next clock tick (to time \( \text{Clk} + 1 \)), so that policies in \( NSF^d \) and \( IS^d \) may change only at clock ticks (and not between).

We call this arrangement a “midnight-shift scheme”, since the underlying idea is the same as that of periodically shifting guards at a museum or a bank. Implementing this scheme is straightforward. For \( \text{R5} \), capabilities carry timestamps. For \( \text{R6} \), administrative operations are executed on an “accumulator” \( \Xi \) instead of \( F \), and at every clock tick, \( F \) is updated to \( \Xi \). Accordingly, we modify the following messages in \( NSF^d \) to obtain the protocol \( NSF^{d+} \).

\[
\begin{align*}
(2) & \quad A \rightarrow U : \text{mac}(\langle U, op, \text{Clk} \rangle, K_{AS}) \quad \text{if} \ F \vdash_U \text{op} \\
(2') & \quad A \rightarrow U : \text{mac}(\langle U, op, \text{Clk} \rangle, K_{AS}) \quad \text{otherwise} \\
(4) & \quad S \rightarrow V : r \quad \text{if} \ \kappa = \text{mac}(\langle \_ , op, \text{Clk} \rangle, K_{AS}) \\
& \quad \quad \quad \text{and} \ \rho[\text{op}] \downarrow \rho'[r] \\
(6) & \quad A \rightarrow W : r \quad \text{if} \ F \vdash_W \theta \text{ and } \Xi[\theta] \downarrow \Xi'[r]\end{align*}
\]

Likewise, we modify the following message in \( IS^d \) to obtain the protocol \( IS^{d+} \).

\[
\begin{align*}
(\text{iv}) & \quad D \rightarrow W : r \quad \text{if} \ F \vdash_W \theta \text{ and } \Xi[\theta] \downarrow \Xi'[r]
\end{align*}
\]

Now a capability that carries \( \text{Clk} \) as its timestamp certifies a particular access decision at the instant \( \text{Clk} \): the meaning is made explicit in the content, which is good practice. However, recall that MACs can be decoded to retrieve their messages. In particular, one can tell the time in \( NSF^{d+} \) by decoding capabilities. Clearly we require that:

**R7** If it is possible to tell the time in \( NSF^{d+} \), it must also be possible to do so in \( IS^{d+} \).

So we must make it possible to tell the time in \( IS^{d+} \). (The alternative is to make it impossible to tell the time in \( NSF^{d+} \), by encrypting the timestamps carried by capabilities. Recall that the notion of “time” here is purely logical.) Accordingly we add the following messages to \( IS^{d+} \).

\[
\begin{align*}
(\text{v}) & \quad U \rightarrow D : () \\
(\text{vi}) & \quad D \rightarrow U : \text{Clk}
\end{align*}
\]

The following result can then be proved (cf. (\*)).

\[ ^2 \text{Some implementation details, such as (R3), are not required for safety.} \]
Theorem 3. \( NS^{d+} \) safely implements \( IS^{d+} \).

This result appears in [11]. Unfortunately, the definition of safety in [11] is rather non-standard. Moreover, beyond this result, security is not considered in [11]. In the rest of this section, we analyze the difficulties that arise for security, and present new results.

It turns out that there are several recipes to break security, and expiry of capabilities is a common ingredient. Clearly, using an expired capability has no counterpart in \( IS^{d+} \). So:

**R8** Any use of an expired capability must block (without any observable effect).

Indeed, security breaks without (R8). Consider the \( NS^{d+} \) pseudo-code below. Informally, stale means “detect any use of an expired capability”. Here \( \kappa \) is a capability for operation \( op \).

\[
t_3 \text{ acquire } \kappa; \text{ use } \kappa; \text{ stale}
\]

Without (R8), \((t3)\) can be distinguished from a \text{false} event by testing the event \text{stale}. But consider implementation instances \( Q_1 \) and \( Q_2 \) with \( op \) as the only operation, whose execution has no observable effect on the store; let \( Q_1 \) run \((t3)\) and \( Q_2 \) run \text{false}. Since \text{stale} cannot be reproduced in the specification, it must map to \text{false}. So the specification instances of \( Q_1 \) and \( Q_2 \) run \text{exec } op; \text{false} and \text{false}. These instances cannot be distinguished.

Moreover, expiry of a capability yields the information that time has elapsed between the acquisition and use of that capability. We may expect that leaking this information is harmless; after all, the elapse of time can be trivially detected by inspecting timestamps. Then why should we care about such a leak? If the adversary knows that the clock has ticked at least once, it also knows that any pending administrative operations have been executed, possibly modifying the access policy. If this information is leaked in a way that cannot be reproduced in the specification, we are in trouble. Any such way allows the adversary to implicitly control the expiry of a capability before its use. (Explicit controls, such as comparison of timestamps, are not problematic, since they can be reproduced in the specification.)

For instance, consider the \( NS^{d+} \) pseudo-code below. Here \( \kappa \) and \( \kappa' \) are capabilities for operations \( op \) and \( op' \), and \( \theta \) modifies access to \( op \).

\[
t_4 \text{ acquire } \kappa';
\begin{align*}
    & \text{ chmod } \theta; \text{ acquire } \kappa; \text{ use } \kappa; \text{ success}; \\
    & \text{ use } \kappa'; \text{ success}
\end{align*}
\]

\[
t_5 \text{ chmod } \theta; \text{ acquire } \kappa; \text{ use } \kappa; \text{ success};
\begin{align*}
    & \text{ acquire } \kappa'; \text{ use } \kappa'; \text{ success}
\end{align*}
\]

Both \((t4)\) and \((t5)\) map to the same \( IS^{d+} \) pseudo-code

\[
\text{ chmod } \theta; \text{ exec } op; \text{ success}; \text{ exec } op'; \text{ success}
\]

But suppose that initially no user has access to \( op \) and all users have access to \( op' \), and \( \theta \) specifies that all users may access \( op \). The intermediate success event is true only
if \( \theta \) is executed; therefore it “forces” time to elapse for progress. Now (t4) and (t5) can be distinguished by testing the final success event. In (t4) \( \kappa' \) must be stale when used, so the event must be false; but in (t5) \( \kappa' \) may be fresh when used, so the event may be true. Therefore, security breaks.

**Security in a special case** One way of plugging such leaks is to consider that the elapse of time is altogether unobservable. (This prospect is not as shocking as it sounds, since here “time” is simply the value of a privately maintained counter.)

We expect that executing an operation has some observable effect. Now if initially a user does not have access to an operation, but that access can be dynamically granted, then the elapse of time can be detected by observing the effect of executing that operation. So we must assume that:

**A2** Accesses cannot be dynamically granted.

On the other hand, we must allow accesses to be dynamically revoked, since otherwise the access policy becomes static. Now if initially a user has access to an operation, but that access can be dynamically revoked, then it is possible to detect the elapse of time if the failure to execute that operation is observable. So we must assume that:

**A3** Any unsuccessful use of a capability blocks (without any observable effect).

Let us now try to adapt the counterexample above with (A2) and (A3). Suppose that initially all users have access to \( op \) and \( op' \), and \( \theta \) specifies that no user may access \( op \). Consider the \( NS^{d+} \) pseudo-code below. Informally, failure means “detect unsuccessful use of a capability”.

\[
\begin{align*}
t6 & \text{ acquire } \kappa'; \\
& \quad \text{ chmod } \theta; \text{ acquire } \kappa; \text{ use } \kappa; \text{ failure}; \\
& \quad \text{ use } \kappa'; \text{ success}
\end{align*}
\]

\[
\begin{align*}
t7 & \text{ chmod } \theta; \text{ acquire } \kappa; \text{ use } \kappa; \text{ failure}; \\
& \quad \text{ acquire } \kappa'; \text{ use } \kappa'; \text{ success}
\end{align*}
\]

Both (t6) and (t7) map to the same \( IS^{d+} \) pseudo-code

\[
\text{ chmod } \theta; \text{ exec } op; \text{ failure; exec } op'; \text{ success}
\]

Fortunately, now (t6) and (t7) cannot be distinguished, since the intermediate failure event cannot be observed if true. (In contrast, recall that the intermediate success event in (t4) and (t5) forces a distinction between them.)

Indeed, with (A2) and (A3) there remains no way to detect the elapse of time, except by comparing timestamps. To prevent the latter, we assume that:

**A4** Timestamps are encrypted.
Let $E_{AS}$ be a secret key shared by $A$ and $S$. The encryption of a term $M$ with $E_{AS}$ under a random coin $m$ is written as $\{m, M\}_{E_{AS}}$. We remove message (4') and modify the following messages in $NS^{d+}$ to obtain the protocol $NS^{d-}$. (Note that randomization takes care of (R4), so capabilities are not required to mention users here.)

(2) $A \rightarrow U : \text{mac}(\langle U, op, \{m, Clk\}_{E_{AS}} \rangle, K_{AS})$

if $F \vdash_U op$

(2') $A \rightarrow U : \text{mac}(\langle U, op, \{m, Clk\}_{E_{AS}} \rangle, \overline{K}_{AS})$

otherwise

(4) $S \rightarrow V : r$ if $\kappa = \text{mac}(\langle \cdot, op, T \rangle, K_{AS})$, $T = \langle \cdot, Clk \rangle_{E_{AS}}$, and $\rho[\cdot] \downarrow \rho'[r]$

Accordingly, we remove the messages (iv'), (v), and (vi) from $IS^{d+}$ to obtain the protocol $IS^{d-}$. We can then prove the following new result (see Section 6):

**Theorem 4.** $NS^{d-}$ securely implements $IS^{d-}$ assuming (A2), (A3), and (A4).

The key observation is that with (A2), (A3), and (A4), time can stand still (so that capabilities never expire).

**Security in the general case** More generally, we may consider plugging problematic leaks by static analysis. (Any such analysis must be incomplete because of the undecidability of the problem.) However, several complications arise in this case.

- The adversary can control the elapse of time by interacting with honest users in subtle ways. Such interactions lead to counterexamples of the same flavor as the one with (t4) and (t5) above, but are difficult to prevent statically without severely restricting the code run by honest users. For instance, even if the suspicious-looking pseudo-code

  \begin{verbatim}
  chmod \theta; acquire \kappa; use \kappa; success
  \end{verbatim}

  is replaced by an innocuous pair of inputs on a public channel $c$, the adversary can still run the same code in parallel and serialize it by a pair of outputs on $c$ (which serve as “begin/end” signals).

- Even if we restrict the code run by honest users, such that every use of a capability can be serialized immediately after its acquisition, the adversary can still force time to elapse after a capability is sent to the file system and before it is examined. Unless we have a way to constrain this elapse of time, we are in trouble.

To see how the adversary can break security by interacting with honest users, consider the $NS^{d+}$ pseudo-code below. Here $\kappa$ is a capability for operation $op$, and $\theta$ modifies access to $op$; further $c()$ and $\overline{w}()$ denote input and output on public channels $c$ and $w$.

\begin{verbatim}
t8 acquire \kappa; use \kappa; c(); chmod \theta; c(); success; \overline{w}()
t9 c(); c(); \overline{w}()
\end{verbatim}

Although $use \kappa$ immediately follows $acquire \kappa$ in (t8), the delay between $use \kappa$ and $success$ can be detected by the adversary to force time to elapse between those events.
Suppose that initially no user has access to $op$ or $op'$, $\theta$ specifies that a honest user $U$ may access $op$, and $\theta'$ specifies that all users may access $op'$. Consider the following context. Here $\kappa'_0$ and $\kappa'_1$ are capabilities for $op'$.

$$\tau();\text{acquire } \kappa'_0;\text{use } \kappa'_0;\text{failure};$$
$$\text{chmod } \theta';\text{acquire } \kappa'_1;\text{use } \kappa'_1;\text{success};\tau()$$

This context forces time to elapse between a pair of outputs on $c$. The context can distinguish $(t8)$ and $(t9)$ by testing output on $w$: in $(t8)$ $\kappa$ does not authorize access to $op$, so $\text{success}$ is false and there is no output on $w$; on the other hand, in $(t9)$ there is. Security breaks as a consequence. Consider implementation instances $Q_1$ and $Q_2$ with $U$ as the only honest user and $op$ and $op'$ as the only operations, such that only $U$ can detect execution of $op$ and all users can detect execution of $op'$; let $Q_1$ run $(t8)$ and $Q_2$ run $(t9)$. The specification instances of $Q_1$ and $Q_2$ run

$$\text{exec } op; c(); \text{chmod } \theta; c(); \text{success}; w();$$
$$c(); w();$$

which cannot be distinguished: the execution of $op$ can always be delayed until $\theta$ is executed, so that $\text{success}$ is true and there is an output on $w$. Intuitively, an execution request in $NS^{d+}$ commits to a time bound (specified by the timestamp of the capability used for the request) within which that request must be processed for progress; but operation requests in $IS^{d+}$ make no such commitment.

To solve this problem, we must assume that:

**A5** In $IS^{d+}$ a time bound is specified for every operation request, so that the request is dropped if it is not processed within that time bound.

Usual (unrestricted) requests now carry a time bound $\infty$. Accordingly we modify the following messages in $IS^{d+}$.

(i) $V \rightarrow D : (op, T)$
(ii) $D \rightarrow V : r$ if $\text{Clk} \leq T$,

$$F \vdash_V \text{op}, \text{and } \rho[\text{op}] \Downarrow \rho'[v]$$

With (A5), using an expired capability now has a counterpart in $IS^{d+}$. Informally, if a capability for an operation $op$ is produced at time $T$ in $NS^{d+}$, then any use of that capability in $NS^{d+}$ maps to an execution request for $op$ in $IS^{d+}$ with time bound $T$. There remains no fundamental difference between $NS^{d+}$ and $IS^{d+}$. We can then prove our main new result (see Section 6):

**Theorem 5** (Main theorem). $NS^{d+}$ securely implements $IS^{d+}$ assuming (A5).\(^3\)

Fortunately, (A5) seems to be a reasonable requirement, and we impose that requirement implicitly in the sequel.

**Discussion** Let us now revisit the principles developed in Sections 2 and 3, and discuss some alternatives.

First recall (R3), where we introduce fake capabilities to prevent premature leaks of information about the access policy $F$. It is reasonable to consider that we do not

\(^3\)This result holds with or without (A5).
care about such leaks, and wish to keep the original message \((2')\) in \(NS^*\). But then we must allow those leaks in the specification. For instance, we can make \(F\) public. More practically, we can add messages to \(IS^*\) that allow a user to know whether it has access to a particular operation.

Next recall \((R5)\) and \((R6)\), where we introduce the midnight-shift scheme. This scheme can be relaxed to allow different capabilities to expire after different intervals, so long as administrative operations that affect their correctness are not executed before those intervals elapse. Let \(\text{delay}\) be a function over users \(U\), operations \(op\), and clock values \(\text{Clk}\) that produces time intervals. We may have that:

\[
\text{R5} \quad \text{Any capability for } U \text{ and } op \text{ that is produced at time } \text{Clk} \text{ expires at time } \text{Clk} + \text{delay}(U, op, \text{Clk}).
\]

\[
\text{R6} \quad \text{If an administrative operation affects the access decision for } U \text{ and } op \text{ and is requested in the interval } \text{Clk}, \ldots, \text{Clk} + \text{delay}(U, op, \text{Clk}) - 1, \text{ it is executed at the clock tick to time } \text{Clk} + \text{delay}(U, op, \text{Clk}).
\]

This scheme remains sound, since any capability for \(U\) and \(op\) that is produced at \(\text{Clk}\) and expires at \(\text{Clk} + \text{delay}(U, op, \text{Clk})\) certifies a correct access decision for \(U\) and \(op\) between \(\text{Clk}, \ldots, \text{Clk} + \text{delay}(U, op, \text{Clk}) - 1\).

Finally, the implementation details in Sections 2 and 3 are far from unique. Guided by the same underlying principles, we can design capabilities in various other ways. For instance, we may have an implementation that does not require \(\text{K}_{AS}\): any capability is of the form \(\text{mac}((\langle U, op, \text{Clk} \rangle, \{m, L\} \text{E}_{AS}), K_{AS})\), where \(m\) is a fresh nonce and \(L\) is the predicate \(F \vdash_U op\). Although this design involves more cryptography than the one in \(NS_{id}^d\), it reflects better practice: the access decision for \(U\) and \(op\) under \(F\) is explicit in the content of any capability that certifies that decision. What does this design buy us? Consider applications where the access decision is not a boolean, but a label, a decision tree, or some arbitrary data structure. The design in \(NS_{id}^d\) requires a different signing key for each value of the access decision. Since the number of such keys may be infinite, verification of capabilities becomes very inefficient. The design above is appropriate for such applications, and we develop it further in Section 7.

### 4 Definitions and proof techniques

Let us now develop formal definitions and proof techniques for security and safety; these serve as background for Section 6 where we present formal models and proofs for security and safety of \(NS_{id}^d\) with respect to \(IS_{id}^d\).

Let \(\preceq\) be a precongruence on processes and \(\simeq\) be the associated congruence. A process \(\Pi\) under a context \(\varphi\) is written as \(\varphi[\Pi]\). Contexts act as tests for behaviors, and \(\Pi \preceq Q\) means that any test that is passed by \(\Pi\) is passed by \(Q\)—in other words, “\(\Pi\) has no more behaviors than \(Q\)”.

We describe an implementation as a binary relation \(\mathcal{R}\) over processes, which relates specification instances to implementation instances. This relation conveniently generalizes a refinement map \([4]\).
Definition 3 (Full abstraction). An implementation $R$ is fully abstract if it satisfies:

\[(\text{Preservation})\]
\[\forall (P, Q) \in R. \forall (P', Q') \in R. \quad P \preceq P' \Rightarrow Q \preceq Q'\]

\[(\text{Reflection})\]
\[\forall (P, Q) \in R. \forall (P', Q') \in R. \quad Q \preceq Q' \Rightarrow P \preceq P'\]

(Preservation) and (Reflection) are respectively soundness and completeness of the implementation under $\preceq$. Security only requires soundness.

Definition 4 (cf. Definition 2 [Security]). An implementation is secure if it satisfies (Preservation).

Intuitively, a secure implementation does not introduce any interesting behaviors—if $(P, Q)$ and $(P', Q')$ are in a secure $R$ and $P$ has no more behaviors than $P'$, then $Q$ has no more behaviors than $Q'$. A fully abstract implementation moreover does not eliminate any interesting behaviors.

Any subset of a secure implementation is secure. Security implies preservation of $\simeq$. Finally, testing itself is trivially secure since $\preceq$ is closed under any context.

Proposition 6. Let $\varphi$ be any context. Then $\{(P, \varphi[P]) \mid P \in W\}$ is secure for any set of processes $W$.

On the other hand, a context may eliminate some interesting behaviors by acting as a test for those behaviors. A fully abstract context does not; it merely translates behaviors.

Definition 5 (Fully abstract context). A context $\varphi$ is fully abstract for a set of processes $W$ if $\{(P, \varphi[P]) \mid P \in W\}$ is fully abstract.

A fully abstract context can be used as a wrapper to account for any benign differences between the implementation and the specification. An implementation is safe if it does not introduce any behaviors modulo such a wrapper.

Definition 6 (cf. Definition 1 [Safety]). An implementation $R$ is safe if there exists a fully abstract context $\phi$ for the set of specification instances such that $R$ satisfies:

\[(\text{Inclusion})\]
\[\forall (P, Q) \in R. \quad Q \preceq \phi[P]\]

Let us see why $\phi$ must be fully abstract in the definition. Suppose that it is not. Then for some $P$ and $P'$, we have $\phi[P] \simeq \phi[P']$ and $P \not\preceq P'$. Intuitively, $\phi$ “covers up” the behaviors of $P$ that are not included in the behaviors of $P'$. Unfortunately, those behaviors may be unsafe. For instance, let $P'$ be a pi calculus process that does not contain public channels, and $\{P'\}$ be the set of specification instances—we consider any output on a public channel to be unsafe. Let $c$ be a public channel; let $P = \langle \tau \rangle$; $P'$ and $\phi = \bullet \mid ! \tau \cdot$. Then $P \not\preceq P'$ and $\phi[P] \simeq \phi[P']$, as required. But clearly $P$ is unsafe by our assumptions; yet $P \preceq \phi[P']$, so that by definition $\{(P', P)\}$ is safe! The definition therefore becomes meaningless.
We now present some proof techniques. A direct proof of security requires mappings between subsets of $\preceq$. Those mappings may be difficult to define and manipulate. Instead a security proof may be built modularly by showing simulations, as in a safety proof. Such a proof requires simpler mappings between processes.

**Proposition 7** (Proof of security). Let $\phi$ and $\psi$ be contexts such that for all $(P, Q) \in R$, $Q \preceq \phi[P]$, $P \preceq \psi[Q]$, and $\phi[\psi[Q]] \preceq Q$. Then $R$ is secure.

**Proof.** Suppose that $(P, Q) \in R$, $P \preceq P'$, and $(P', Q') \in R$. Then $Q \preceq \phi[P] \preceq \phi[P'] \preceq \phi[\psi[Q']] \preceq Q'$.

Intuitively, $R$ is secure if $R$ and $R^{-1}$ both satisfy (INCLUSION), and the witnessing contexts “cancel” each other. A simple technique for proving full abstraction for contexts follows as a corollary.

**Corollary 8** (Proof of full abstraction for contexts). Let there be a context $\varphi^{-1}$ such that for all $P \in W$, $\varphi^{-1}[\varphi[P]] \simeq P$. Then $\varphi$ is a fully abstract context for $W$.

**Proof.** Take $\varphi = \varphi^{-1}$ and $\psi = \varphi$ in the proposition above to show that $\{(\varphi[P], P) \mid P \in W\}$ is secure. The converse follows by Proposition.

**Theory for the applied pi calculus** Let $a, b, \ldots$ range over names, $u, v, \ldots$ over names and variables, $M, N, \ldots$ over terms, and $A, B, \ldots$ over extended processes. Semantic relations include the binary relations $\equiv$, $\rightarrow$, and $\stackrel{\ell}{\rightarrow}$ over extended processes (structural equivalence, reduction, and labeled transition); here labels $\ell$ are of the form $a(M)$ or $(\nu \bar{a}) \pi(\bar{v})$ (where $a \notin \bar{u}$ and $\bar{a} \subseteq \bar{v}$). Both $\rightarrow$ and $\rightarrow^*$ are closed under $\equiv$ and $\rightarrow$ is closed under arbitrary evaluation contexts.

We recall some theory on may testing for applied pi calculus programs.

**Definition 7** (Barb). A barb $\downarrow a$ is a predicate that tests possible output on $a$; we write $A \downarrow a$ if $A \xrightarrow{(\nu \bar{a}) \pi(\bar{v})} B$ for some $B$, $\bar{v}$, and $\bar{u}$. A weak barb $\downarrow^w a$ tests possible eventual output on $a$, i.e., $\downarrow^w a \triangleq \rightarrow^* \downarrow a$.

**Definition 8** (Frame). Let $A$ be closed. Then we have $A \equiv (\nu \bar{a})(\sigma \mid P)$ for some $\bar{a}$, $\sigma$, and $P$ such that $\text{fv(rng}(\sigma)) \cup \text{fv}(P) = \emptyset$; define $\text{frame}(A) \equiv (\nu \bar{a}) \sigma$.

**Definition 9** (Static equivalence). Let $A$ and $B$ be closed. Then $A$ is statically equivalent to $B$, written $A \approx_a B$, if there exists $\bar{a}$, $\sigma$, and $\sigma'$ such that $\text{frame}(A) \equiv (\nu \bar{a}) \sigma$, $\text{frame}(B) \equiv (\nu \bar{a}) \sigma'$, $\text{dom}(\sigma) \equiv \text{dom}(\sigma')$, and for all $M$ and $N$,

$$\{\bar{a}\} \cap (\text{fn}(M) \cup \text{fn}(N)) = \emptyset \implies M \sigma = N \sigma \iff M \sigma' = N \sigma'.$$

**Proposition 9.** $A \approx_a B$ if and only if $\text{frame}(A) \simeq \text{frame}(B)$.

**Proof.** By induction on the structure of closing evaluation contexts.

We can prove $\preceq$ by showing a simulation relation that approximates $\leq$.  

14
**Definition 10** (Simulation preorder). Let \( \preceq \) be the largest relation \( S \) such that for all \( A \) and \( B \), \( (A, B) \in S \) implies

- \( A \approx_s B \)
- \( \forall A'. A \rightarrow A' \Rightarrow \exists B'. B \rightarrow^* B' \land (A', B') \in S \)
- \( \forall A', \alpha. A \xrightarrow{\ell} A' \Rightarrow \exists B'. B \rightarrow^* \xrightarrow{\ell} B' \land (A', B') \in S \)

**Proposition 10** (Proof of testing precongruence). \( \preceq \subseteq \leq \).

## 5 Models and proofs for static access policies

We now present implementation and specification models and security proofs for static access policies. Models and proofs for dynamic access policies follow essentially the same routine, and are presented in the next section.

### 5.1 Preliminaries

We fix an equational theory \( \Sigma \) with the following properties.

- \( \Sigma \) includes a theory of natural numbers with symbols \( 0 \) (zero), \( +1 \) (successor), and \( \leq \) (less than or equal to).
- \( \Sigma \) includes a theory of finite tuples with symbols \( \langle \_ \_ \rangle \) (indexed concatenate) and \( \_ \_ \_ \) (indexed project).
- \( \Sigma \) contains exactly one equation that involves the symbol \( \text{mac} \), which is
  \[ \text{msg}(\text{mac}(x, y)) = x \]

Clients are identified by natural numbers; we fix a finite subset \( I \) of \( \mathbb{N} \) and consider any user not identified in \( I \) to be dishonest.

File-system code and other processes are conveniently modeled by parameterized process expressions, whose semantics are defined (recursively) by extending the usual semantic relations \( \equiv, \rightarrow, \) and \( \xrightarrow{\ell} \).

### 5.2 Models

Figures 1 and 2 show applied pi calculus models for the file systems under study. We ignore the rules in the inner boxes in these figures (labeled (DUMMY...)) in a first reading.

Figure 1 models a traditional file system (with local access control). The file system is parameterized by an access policy \( F \), a store \( \rho \), and a renaming \( \eta \) of its default interface. That interface includes a channel \( \beta^\rho_k \) for every \( k \in \mathbb{N} \); intuitively, a user identified by \( k \) may send operation requests on this channel.

Processes \( \text{Req}_k(F, op, n) \) and \( \text{EOk}(M, op, n) \) denote internal states. In the equational theory \( \text{auth}(F, k, op) = \text{ok} \) means that user \( k \) may access \( op \) under \( F \), and
\[ k \in \mathbb{N} \]
\[ \text{TFS}(F, \rho)^{\eta} \equiv \eta(\beta_k)(\text{op}, x); \text{Req}_k(F, \text{op}, x)^{\eta} | \text{TFS}(F, \rho)^{\eta} \]

\[ \text{OP} \]
\[ \text{perm}(F, k, \text{op}) = L \]
\[ \text{Req}_k(F, \text{op}, M) \rightarrow \text{EOk}(L, \text{op}, M) \]

\[ \text{EXEC} \]
\[ \text{exec}(L, \text{op}, \rho) = \langle N, \rho' \rangle \]
\[ \text{EOk}(L, \text{op}, M) | \text{TFS}(F, \rho)^{\eta} \rightarrow \bar{M}\langle N \rangle | \text{TFS}(F, \rho')^{\eta} \]

\[ \text{DUMMY AUTH} \]
\[ j \in \mathbb{N} \setminus \mathcal{I} \]
\[ \tau^{\text{TS}}_{\mathcal{I}} \equiv \eta(\alpha_j)(\text{op}, x); \overline{\text{mac}}((j, \text{op}, K_\gamma)) | \tau^{\text{TS}}_{\mathcal{I}} \]

\[ \text{DUMMY EXEC} \]
\[ j \in \mathbb{N} \setminus \mathcal{I} \]
\[ \tau^{\text{TS}}_{\mathcal{I}} \equiv \eta(\beta_j)(\kappa, x); \text{DReq}(\kappa, x)^{\eta} | \tau^{\text{TS}}_{\mathcal{I}} \]

\[ \text{DUMMY CAP} \]
\[ \kappa = \text{mac}(\text{msg}(\kappa), K_\gamma) \]
\[ \text{msg}(\kappa) = (j, \text{op}) \]
\[ j \in \mathbb{N} \setminus \mathcal{I} \]
\[ \text{DReq}(\kappa, M)^{\eta} \rightarrow \eta(\beta_j^\kappa)(\text{op}, M) \]

\[ \text{EXEC} \]
\[ \text{CReq}_k(F, \text{op}, x)^{\eta} | \text{NAFS}(F, \rho)^{\eta} \]
\[ \text{NAFS}(F, \rho)^{\eta} \equiv \eta(\alpha_k)(\text{op}, x); \text{CReq}_k(F, \text{op}, x)^{\eta} | \text{NAFS}(F, \rho)^{\eta} \]

\[ \text{OP} \]
\[ \text{verif}(\kappa) = L \quad L \in \{\text{true, false}\} \]
\[ \text{msg}(\kappa) = (\omega, \text{op}) \]
\[ \text{Req}(\kappa, M) \rightarrow \text{EOk}(L, \text{op}, M) \]

\[ \text{DUMMY OP} \]
\[ j \in \mathbb{N} \setminus \mathcal{I} \]
\[ \tau^{\text{TS}}_{\mathcal{I}} \equiv \eta(\beta_j^\kappa)(\kappa, x); \text{DReq}(\kappa, x)^{\eta} | \tau^{\text{TS}}_{\mathcal{I}} \]

\[ \text{DUMMY AUTH & EXEC} \]
\[ \text{DReq}(\text{op}, M)^{\eta} \equiv (\forall c \rightarrow \eta(\alpha_j)\langle \text{op, c}; c(\kappa); \eta(\beta_j^\kappa)\rangle(k, M) \]

Figure 1: A traditional file system with local access control

Figure 2: A network-attached file system with distributed access control
exec\((L, op, \rho) = \langle N, \rho' \rangle\) means that the execution of \(op\) on store \(\rho\) under decision \(L\) returns \(N\) and store \(\rho'\). Decisions are derived by \(\text{perm\((\_, \_, \_\))}\) as follows.

\[
L = \text{true if } \text{auth}(F, k, op) = \text{ok}, = \text{false otherwise}
\]

\[
\text{perm}(F, k, op) \triangleq L
\]

A traditional storage system may be described as

\[
(\nu_{i \in \mathbb{I}} \beta_i^\rho)(C \mid \text{IFS}(F, \rho))
\]

Here \(C\) is code run by honest users; the file-system exports the default interface (implicitly renamed by “identity”), and channels associated with honest users are hidden from the context. The context may be arbitrary and is left implicit; in particular, channels associated with dishonest users are available to the context.

Figure 2 models a network-attached file system (with distributed access control). As above, the file system is parameterized by an access policy \(F\), a store \(\rho\), and a renaming \(\eta\) of its default interface. That interface includes channels \(\alpha_k\) and \(\beta_k\) for every \(k \in \mathbb{N}\); intuitively, a user identified by \(k\) may send authorization requests on \(\alpha_k\) and execution requests on \(\beta_k\).

Processes \(\text{CReq}_k(F, op, c)\), \(\text{Req}(\kappa, n)\), and \(\text{EOk}(M, op, n)\) denote internal states. In the equational theory \(\text{auth}(F, k, op) = \text{ok}\) and \(\text{exec}(L, op, \rho) = \langle N, \rho' \rangle\) have the same meanings as above. Capabilities and decisions are derived by \(\text{cert\((\_, \_, \_\))}\) and \(\text{verif\((\_\))}\) as follows.

\[
a = K_{MD} \text{ if } \text{auth}(F, k, op) = \text{ok}, = K'_M \text{ otherwise}
\]

\[
\text{cert}(F, k, op) \triangleq \text{mac}((k, op), a)
\]

\[
L = \text{true if } \kappa = \text{mac}(\text{msg}(\kappa), K_{MD}), = \text{false if } \kappa = \text{mac}(\text{msg}(\kappa), K'_M)
\]

\[
\text{verif}(\kappa) \triangleq L
\]

A network-attached storage system may be described as

\[
(\nu_{i \in \mathbb{I}} \alpha_i^\rho)(C \mid (\nu K_{MD} K'_M) \text{NAFS}(F, \rho))
\]

As above, \(C\) is code run by honest users; the file-system exports the default interface and hides the keys that authenticate capabilities. Channels associated with honest users are hidden from the context. The context may be arbitrary and is left implicit; in particular, channels associated with dishonest users are available to the context.

### 5.3 Proofs of security

We prove that the implementation is secure, safe, and fully abstract with respect to the specification. We begin by outlining the proofs, and then present details.

#### 5.3.1 Outline

Let \(F\), \(\rho\), and \(C\) range over access policies, stores, and code for honest users that are “wellformed” in the implementation. Let \([\_]\) abstract such \(F\), \(\rho\), and \(C\) in the
specification. We define
\[
\mathcal{R} = \bigcup_{F, \rho, C} \{ (\nu \in I [\beta_i]) ([C] | TFS([F], [\rho])) , (\nu \in I [\alpha_i \beta_i]) ([C] | (\nu K_{MD} K_M') NAFS(F, \rho)) \}
\]
We prove that \( \mathcal{R} \) is secure by showing contexts \( \phi \) and \( \psi \) such that:

**Lemma 11.** For any \( F, \rho, \) and \( C, \)

1. \( (\nu \in I [\beta_i]) ([C] | (\nu K_{MD} K_M') NAFS(F, \rho)) \preceq \phi((\nu \in I [\beta_i]) ([C] | TFS([F], [\rho]))) \)
2. \( (\nu \in I [\beta_i]) ([C] | TFS([F], [\rho])) \preceq \psi((\nu \in I [\alpha_i \beta_i]) ([C] | (\nu K_{MD} K_M') NAFS(F, \rho))) \)
3. \( \phi[\psi([\nu \in I [\beta_i]) ([C] | (\nu K_{MD} K_M') NAFS(F, \rho)))] \preceq (\nu \in I [\alpha_i \beta_i]) ([C] | (\nu K_{MD} K_M') NAFS(F, \rho)) \)

Proposition\(^7\) then applies. Moreover we show:

**Lemma 12.** For any \( F, \rho, \) and \( C, \)

\[
\psi[\phi((\nu \in I [\beta_i]) ([C] | TFS([F], [\rho])))] \preceq (\nu \in I [\beta_i]) ([C] | TFS([F], [\rho]))
\]

Now \( \mathcal{R}^{-1} \) is secure by Proposition\(^7\) Thus \( \mathcal{R} \) is proved fully abstract. Moreover Lemmas\(^11\)–2 already imply the converse of Lemma\(^12\) so \( \phi \) is a fully abstract context by Corollary\(^8\) (taking \( \phi^{-1} = \psi \)). Thus \( \mathcal{R} \) is proved safe.

We now revisit Figures\(^1\) and\(^2\) and focus on the rules in the inner boxes. Those rules define processes \( \triangleright_{TSA} \) and \( \triangleright_{NAS} \). Intuitively, these processes translate public requests from \( NAS^a \) to \( TSA^a \) and from \( TSA^a \) to \( NAS^a \). Let \( \mathring{\alpha}_{TSA} \) and \( \mathring{\alpha}_{NAS} \) include the public interfaces of \( TSA^a \) and \( NAS^a \). We define

\[
\phi = (\nu \mathring{\alpha}_{TSA}) (\bullet | \triangleright_{TSA}) \quad \psi = (\nu \mathring{\alpha}_{NAS}) (\bullet | \triangleright_{NAS})
\]
The abstraction function $\llbracket \cdot \rrbracket$ is shown in Figure 3. Here $\mathcal{A}$ contains special names whose uses in well-formed code are either disciplined or forbidden.

$$\mathcal{A} \triangleq \left\{ \alpha_i, \beta_i \mid i \in \mathcal{I} \right\} \cup \left\{ \alpha_j, \beta_j, \beta_j' \mid j \in \mathbb{N} \setminus \mathcal{I} \right\} \cup \left\{ K_{MD}, K'_M, K_? \right\}$$

The names in $\{\alpha_j, \beta_j, \beta_j' \mid j \in \mathbb{N} \setminus \mathcal{I}\} \cup \{K_?\}$ are invented to simplify proofs below.

### 5.3.2 Simulation relations

Figures 4, 5, and 6 show simulation relations for Lemma 11.1–3. All these relations are closed under $\llbracket \cdot \rrbracket$. Here $\eta_1$ and $\eta_2$ rename the public interfaces of $NAS^A$ and $TS^A$ and $\eta_3$ renames the private authentication keys $K_{MD}$ and $K'_M$.

$$\begin{align*}
\eta_1 & \triangleq \left[ \alpha_j \mapsto \alpha_j, \beta_j \mapsto \beta_j \mid j \in \mathbb{N} \setminus \mathcal{I} \right] \\
\eta_2 & \triangleq \left[ \beta_j' \mapsto \beta_j' \mid j \in \mathbb{N} \setminus \mathcal{I} \right] \\
\eta_3 & \triangleq \left[ a \mapsto K_? \mid a \in \{K_{MD}, K'_M\} \right]
\end{align*}$$

These renamings map to names in $\mathcal{A}$ that do not occur in well-formed code (see Figure 4). In particular, the purpose of $\eta_1$ and $\eta_2$ is to rename some public channels to fresh ones that can be hidden by restriction in $\psi$ and $\phi$. (A similar purpose is served by...
for every execution request.

Finally, by $S$ a network-attached storage system may be simulated by a traditional storage system by forwarding public requests directed at NAFS to a hidden TFS interface (via $\phi$). Symmetrically, by $T$ a traditional storage system may be simulated by a network-attached storage system by forwarding public requests directed at TFS to a hidden NAFS interface (via $\psi$). Finally, by $U$ a network-attached storage system may simulate another network-attached storage system by filtering requests directed at NAFS through a hidden TFS interface before forwarding them to a hidden NAFS interface (via $\phi[\psi]$). This rather mysterious detour forces a fresh capability to be acquired for every execution request.

---

**Figure 5:** Simulation relation for Lemma 11.2 ($\preceq \psi[\_]$)

quantification in logic.) Hiding those names strengthens Lemmas 11.1–2 while not affecting their proofs; but more importantly, the restrictions are required to prove Lemma 11.3. Further the purpose of $\eta_3$ is to abstract terms that may be available to contexts. Such terms must be of type Export; intuitively, $K_{MD}$ and $K_M'$ may appear only as authentication keys in capabilities issued to dishonest users.

$$\forall x \in \text{dom}(\sigma) \Rightarrow \exists i \in I, \op, \Gamma(x) = \text{Cert}(i, \op) \land \sigma(x) = \text{cert}(F, i, \op)$$

We show that term abstraction preserves equivalence in the equational theory.

**Lemma 13.** Suppose that $M : F$ Export and $N : F$ Export. Then $M = N$ iff $\eta_3(M) = \eta_3(N)$.

This lemma is required to show static equivalence in proofs of soundness for the relations $S$, $T$, and $U$ in Figures 4, 5, and 6, which in turn lead to Lemma 11. We prove that those relations are included in the simulation preorder.

**Lemma 14.** $S \subseteq \preceq$, $T \subseteq \preceq$, and $U \subseteq \preceq$. 
Figure 6: Simulation relation for Lemma 11.3 (\(\phi[\psi[\phi]] \leftrightarrow \phi\))
Lemma 15. \( V \subseteq \subseteq \). Thus \( \mathcal{R} \) is safe and fully abstract.

By definition of \( \subseteq \) and alphaconversion to default public interfaces, we have for any \( F, \rho, \text{ and } C \):

1. \((\nu_{i \in I} \alpha_i \beta_i)(C \mid (\nu K_{MD} K_{M}')\mathrm{NAFS}(F, \rho)) \not\approx (\nu_{i \in I} \beta_i')(\phi(C \mid \mathrm{TFS}(F', \rho]))\)

2. \((\nu_{i \in I} \beta_i')(C \mid \mathrm{TFS}(F', \rho)) \approx \phi(C \mid \mathrm{TFS}(F', \rho))\)

3. \(\phi(C \mid (\nu K_{MD} K_{M}')\mathrm{NAFS}(F, \rho)) \not\approx (\nu_{i \in I} \alpha_i \beta_i)(C \mid (\nu K_{MD} K_{M}')\mathrm{NAFS}(F, \rho)))\)

Lemma 11 follows by Proposition 10. Thus \( \mathcal{R} \) is secure.

Further, Figure 7 shows a simulation relation for Lemma 12. We prove that the relation \( V \) is included in the simulation preorder.

Lemma 15. \( V \subseteq \subseteq \). Thus \( \mathcal{R} \) is safe and fully abstract.
6 Models and proofs for dynamic access policies

Next we present models and proofs for dynamic access policies, following the routine of Section 5.

6.1 Models

The models extend those in Section 5 and are shown in Figures 8 and 9. (As usual, we ignore the rules in the inner boxes in a first reading.) Interfaces are extended with channels $\delta_k$ and $\delta_k^\circ$ for every $k$, on which users identified by $k$ send administration requests in the implementation and the specification.

In the equational theory $\text{auth}(F, k, op) = \text{ok}$ and $\text{exec}(L, op, \rho) = \langle N, \rho' \rangle$ have the same meanings as in Section 5. Capabilities are derived by $\text{cert}(\_ \_ \_ \_ \_ \_)$ as follows.

$\begin{align*}
a & = K_{MD} \text{ if } \text{auth}(F, k, op) = \text{ok}, \\
\text{cert}(F, k, op, Clk) & \triangleq \text{mac}(\langle k, op, Clk \rangle, a)
\end{align*}$

Recall that administrative operations scheduled at time $Clk$ are executed at the next clock tick (to $Clk + 1$). In the equational theory $\text{push}(L, adm, \Xi, Clk) = \langle N, \Xi' \rangle$ means that an administrative operation $adm$ pushed on schedule $\Xi$ under decision $L$ at $Clk$ returns $N$ and the schedule $\Xi'$; and $\text{sync}(F, \Xi, Clk) = F'$ means that an access policy $F$ synchronized under schedule $\Xi$ at $Clk$ returns the access policy $F'$.

A traditional storage system may be described as

$(\nu_i \in T \alpha_i \beta_i \delta_i)(C | \text{TFS}(F, \emptyset, 0, \rho))$

where $C$ is code run by honest users, $F$ is an access policy and $\rho$ is a store; initially the schedule is empty and the time is 0.

Similarly a network-attached storage system may be described as

$(\nu_i \in T \alpha_i \beta_i \delta_i)(C | (\nu K_{MD} K'_{MD}) \text{NAFS}(F, \emptyset, 0, \rho))$

As usual, let $F$, $\rho$, and $C$ range over access policies, stores, and code for honest users that are “wellformed” in the implementation, and let $\lceil \_ \rceil$ abstract such $F$, $\rho$, and $C$ in the specification. We define

$R = \bigcup_{F, \rho, C} \{ (\nu_i \in T \alpha_i \beta_i \delta_i)(C | \text{TFS}([F], \emptyset, 0, [\rho])) \\
(\nu_i \in T \alpha_i \beta_i \delta_i)(C | (\nu K_{MD} K'_{MD}) \text{NAFS}(F, \emptyset, 0, \rho)) \}$

Figure 10 shows the abstraction function $\lceil \_ \rceil$. Here

$A = \{ \alpha_j, \beta_j, \delta_j, \alpha_j^\circ, \beta_j^\circ, \delta_j^\circ | j \in N \setminus I \} \cup \{ K_{MD}, K'_{MD}, K \} \cup \{ \alpha_i, \beta_i, \delta_i | i \in I \}$

6.2 Examples of security

At this point we revisit the “counterexamples” in Section 3. By modeling them formally in this setting, we show that those counterexamples are eliminated.

Recall (1) and (2).
Figure 8: A traditional file system with local access control
Figure 9: A network-attached file system with distributed access control
\[
\begin{align*}
\text{fn}(M) \cap (A \cup \{\alpha_j, \beta_j, \delta_j \mid j \in \mathbb{N}\}) &= \emptyset & [P]_\Gamma &= Q && \Gamma \supseteq \{\alpha_j, \beta_j, \delta_j \mid j \in \mathbb{N}\} \\
[M] &= M & [P]_\Gamma &= Q
\end{align*}
\]

\[
\begin{align*}
\vdots \quad i \in I & \quad \text{fn}(\text{adm}, M) \cap \text{dom}(\Gamma) = \emptyset \\
& \quad \delta_i(\text{adm}, M); P_{\Gamma} = \delta_i^T(\text{adm}, M); [P]_{\Gamma} \\
\end{align*}
\]

\[
\begin{align*}
\vdots \quad i \in I & \quad \text{fn}(c, x) \cap \text{dom}(\Gamma) = \emptyset \\
& \quad c \notin \text{fn}(P) \\
\end{align*}
\]

\[
\begin{align*}
\{i, i'\} \subseteq I & \quad \Gamma(x) = \text{Cert}(i', \text{op}) \quad \text{fn}(\text{op}, M) \cap \text{dom}(\Gamma) = \emptyset \\
& \quad \delta_i(x, M); P_{\Gamma} = \delta_i^T(\text{op}, x, M); [P]_{\Gamma} \\
\end{align*}
\]

Figure 10: Abstraction function

t1 acquire \(\kappa\); chmod \(\zeta\); use \(\kappa\); success \(\kappa\)

t2 chmod \(\zeta\); acquire \(\kappa\); use \(\kappa\); success \(\kappa\)

The following fragments of \(NAS^d\) code formalize these traces.

\begin{enumerate}
\item\(I_1\) \((vc) \; \overline{\delta_i}(op, c); c(\kappa); (vm) \; \overline{\delta_i}(\zeta, m); m(z); (vm) \; \overline{\beta_i}(\kappa, n); n(x); [\text{success}(x)] \; \overline{\text{adm}}(\cdot)\)
\item\(I_2\) \((vm) \; \overline{\delta_i}(\zeta, m); m(z); (vc) \; \overline{\delta_i}(op, c); c(\kappa); (vm) \; \overline{\beta_i}(\kappa, n); n(x); [\text{success}(x)] \; \overline{\text{adm}}(\cdot)\)
\end{enumerate}

This code is abstracted to the following fragments of \(TS^d\) code.

\begin{enumerate}
\item\(S_1\) \((vc) \; \overline{\alpha_i}(c); c(\tau); (vm) \; \overline{\alpha_i}(\zeta, m); m(z); (vm) \; \overline{\beta_i}(op, \tau, n); n(x); [\text{success}(x)] \; \overline{\text{adm}}(\cdot)\)
\item\(S_2\) \((vm) \; \overline{\delta_i}(\zeta, m); m(z); (vc) \; \overline{\delta_i}(c); c(\tau); (vm) \; \overline{\beta_i}(op, \tau, n); n(x); [\text{success}(x)] \; \overline{\text{adm}}(\cdot)\)
\end{enumerate}

Now whenever \(I_1\) and \(I_2\) can be distinguished, so can \(S_1\) and \(S_2\). Indeed the time bound \(\tau\) is the same as the timestamp in \(\kappa\); so (in particular) the operation request in \(S_1\) is dropped whenever the execution request in \(T_1\) is dropped.

A similar argument counters the “dangerous” example with \(I_4\) and \(I_5\):

\begin{enumerate}
\item\(I_4\) acquire \(\kappa\); chmod \(\zeta\); acquire \(\kappa'\); use \(\kappa'\); success \(\kappa'\); use \(\kappa\); success \(\kappa\)
\item\(I_5\) chmod \(\zeta\); acquire \(\kappa'\); use \(\kappa'\); success \(\kappa'\); acquire \(\kappa\); use \(\kappa\); success \(\kappa\)
\end{enumerate}

Finally, recall \((I_8)\) and \((I_9)\):

\begin{enumerate}
\item\(I_8\) acquire \(\kappa\); use \(\kappa\); \(c()\); chmod \(\zeta\); \(c()\); success \(\kappa\); \(\overline{\text{adm}}(\cdot)\)
\item\(I_9\) \(c(); c(); \overline{\text{adm}}(\cdot)\)
\end{enumerate}

The following fragment of \(NAS^d\) code formalizes \((I_8)\).

\[
\begin{align*}
(I_8) \quad \text{vm}(\zeta, m); m(\kappa); (vm) \; \overline{\beta_i}(\kappa, n); \\
& \quad c(); (vm) \; \overline{\delta_i}(\zeta, m); m(z); c(); n(x); [\text{success}(x)] \; \overline{\text{adm}}(\cdot)
\end{align*}
\]
This code is abstracted to the following fragment of $TS^d$ code.

$$S3 \quad (vm) \quad \alpha^i_m \langle m \rangle \colon m(\tau); (vm) \quad \beta^i \langle op, \tau, n \rangle; c() ; (vm) \quad \delta^i \langle \zeta, m \rangle ; m(z) ; c() ; n(x) ; \{success(x) \} \nu()$$

A $NAS^d$ context distinguishes (I3) and (t9):

$$\nu() \colon \alpha^j \langle op', m_0 \rangle ; m_0(\kappa'_0) ; \beta^j \langle \kappa'_0, n_0 \rangle ; n_0(x) ; \{failure(x) \}$$

$$\delta^j \langle \zeta, p \rangle \colon \alpha^j \langle op', m_1 \rangle ; m_1(\kappa'_1) ; \beta^j \langle \kappa'_1, n_1 \rangle ; n_1(x) ; \{success(x) \} \nu()$$

But likewise a $TS^d$ context distinguishes (S3) and (t9):

$$\nu() \colon \alpha^j \langle m_0 \rangle ; m_0(\tau'_0) ; \beta^j \langle op', \tau'_0, n_0 \rangle ; n_0(x) ; \{failure(x) \}$$

$$\delta^j \langle \zeta, p \rangle \colon \alpha^j \langle m_1 \rangle ; m_1(\tau'_1) ; \beta^j \langle op', \tau'_1, n_1 \rangle ; n_1(x) ; \{success(x) \} \nu()$$

### 6.3 Proofs of security

We show that $R$ is secure, safe, and fully abstract. Recall the contexts $\phi$ and $\psi$ defined in Section 5. The processes $\gamma_{NAS}^d$ and $\gamma_{TS}^d$ are redefined in the inner boxes in Figures 8 and 9. In particular, the rule (DUMMY OP REQ) in Figure 9 translates time-bounded operation requests by $TS^d$ contexts.

Simulation relations for security are shown in Figures 11, 12, and 13 and a simulation relation for safety and full abstraction is shown in Figure 14. Here

$$\eta_1 \triangleq \{ \alpha_j \rightarrow \alpha_{j'}, \beta_j \rightarrow \beta_{j'}, \delta_j \rightarrow \delta_{j'} | j \in \mathbb{N} \}$$

$$\eta_2 \triangleq \{ \alpha^o_j \rightarrow \alpha^o_{j'}, \beta^o_j \rightarrow \beta^o_{j'}, \delta^o_j \rightarrow \delta^o_{j'} | j \in \mathbb{N} \}$$

A binary relation $\omega \vdash \omega$ ("leads-to") is defined over the product of access policies and clocks. Access policies may change at clock ticks (but not between).

$$F', \text{Clk}' \vdash F, \text{Clk} \triangleq (\text{Clk}' \prec \text{Clk}) \lor (\text{Clk}' = \text{Clk} \land F' = F)$$

As usual, any term that may be available to contexts must be of type $\text{Export}$.

$$N = N' \sigma \quad \{ K_{MD, K_M}, K_I \} \cap \text{fn}(N') = \emptyset \quad \forall \nu \in \text{rng}(\sigma). \exists j \in \mathbb{N}, op, \text{Clk}' \quad \text{op} : \text{F, Clk Export} \quad \land \quad (\text{F(\text{Clk}')}, \text{Clk}' \vdash F, \text{Clk}) \quad \land \quad \text{L = cert(F(\text{Clk}'), j, op, \text{Clk}')}$$

$$N : \text{F, Clk Export}$$

We prove that the relations $S$, $T$, and $U$ in Figures 11, 12, and 13 are included in the simulation preorder. Some interesting points in those proofs are listed below.

- In Section 5 when an operation request is sent in $TS^s$ we send an appropriate authorization request in $NAS^s$, obtain a capability, and send an execution request with that capability (see $T$ in Figure 5). In contrast, here when an operation request is sent in $TS^d$ we wait after sending an appropriate authorization request
Figure 11: Simulation relation for Lemma 16.1 (\(\preceq_{\phi}\))
∀x. x ∈ X \implies \exists \beta', \text{Clk}' \sim F, \text{Clk} \land \sigma'(x) = \text{Clk}' \\
\land \Gamma(x) = \text{Cert}(i, \sigma) \land \sigma'(x) = \text{cert}(F', i, \sigma')

\begin{array}{c}
i \in I \quad \text{fn}(\sigma, i, \sigma) \land A = \emptyset \\
\text{TReq}(\sigma, i, \sigma) \quad T^\text{R}_{\sigma, i, \sigma} Q \quad \Gamma(x) = \text{Cert}(i, \sigma)
\end{array}

\begin{array}{c}
P \quad T^\text{T}_{\sigma, i, \sigma} Q' \\
\nu \in \mathcal{L} \quad P' \quad T^\text{T}_{\nu \in \mathcal{L} P'} Q' \\
(\nu \in \mathcal{L} \sigma, \beta, \delta) (P \quad P' \quad \Pi_1 \in \mathcal{L} P' \quad T^\text{T}_{\nu \in \mathcal{L} \sigma, \beta, \delta} (\nu K_D K_M)(Q | Q') \quad \Pi_1 \in \mathcal{L} Q')
\end{array}

\text{(SYSTEM CODE)}

\begin{array}{c}
P \quad T' Q \\
(\nu \in \mathcal{L} \sigma | P) \quad T (\nu \in \mathcal{L} \sigma | (\nu \in \mathcal{L} \sigma, \beta, \delta) (Q | Q') \quad \Pi_1 \in \mathcal{L} Q')
\end{array}

Figure 12: Simulation relation for Lemma[16]2 (\ Bene_\psi[2])
Figure 13: Simulation relation for Lemma 16.3
\[
\begin{align*}
\text{File Systems} & : \forall r \in \mathcal{L}, P_r, \mathcal{V}_r, \text{fs}(\Xi, \rho) \cap A = \emptyset \\
\text{Honest Users} & : \forall x, i \in \text{dom}(\sigma) \implies \exists \text{Clk}', i \in I, \text{Clk}' \leq \text{Clk} \land I(x) = \text{Cert}(i, \text{op}) \land \sigma(x) = \text{Clk}'
\end{align*}
\]
in NAS$^d$ (see $T$ in Figure 12); we continue only when that operation request in $TS^d$ is processed, when we obtain a capability in NAS$^d$, send an execution request with that capability, and process the execution request.

But why wait? Suppose that the operation request in $TS^d$ carries a time bound $\infty$; now if we obtain a capability in NAS$^d$ before the operation request in $TS^d$ is processed, we commit to a finite time bound, which breaks the simulation.

- As before, $\phi[\psi]$ forces a fresh capability to be acquired for every execution request by filtering execution requests in NAS$^d$ through $TS^d$ and back. When an execution request is sent in NAS$^d$ under $\phi[\psi]$ we send an execution request with the same capability in NAS$^d$ (see $U$ in Figure 13). But under $\phi[\psi]$ a fresh capability is obtained and the execution request is sent again with the fresh capability. If the capability in the original request expires before the fresh capability, the simulation breaks. Fortunately operation requests in $TS^d$ carry time bounds, so we can communicate this expiry bound through $TS^d$. In fact there seems to be no way around this problem unless time bounds can be specified in operation requests in $TS^d$!

By Proposition 10 we have:

**Lemma 16.** For any $F$, $\rho$, and $C$,

1. $(\nu_{\iota_1}^{\alpha_{i_1}}{\delta_{i_1}}) (C \mid (\nu K_M D K'_M) NAFS(F, \emptyset, 0, \rho))$
   \[\preceq\phi[(\nu_{\iota_1}^{\alpha_{i_1}}{\delta_{i_1}}) (C \mid \text{TFS}([F], \emptyset, 0, [\rho]))] \]

2. $(\nu_{\iota_2}^{\alpha_{i_2}}{\delta_{i_2}}) \in TFS([F], \emptyset, 0, [\rho])$
   \[\preceq\psi[(\nu_{\iota_2}^{\alpha_{i_2}}{\delta_{i_2}}) (C \mid (\nu K_M D K'_M) NAFS(F, \emptyset, 0, \rho))] \]

3. $\phi[\psi[(\nu_{\iota_3}^{\alpha_{i_3}}{\delta_{i_3}}) (C \mid (\nu K_M D K'_M) NAFS(F, \emptyset, 0, \rho))]]$
   \[\preceq(\nu_{\iota_3}^{\alpha_{i_3}}{\delta_{i_3}}) (C \mid (\nu K_M D K'_M) NAFS(F, \emptyset, 0, \rho)) \]

So by Proposition 7, $R$ is secure.

Further we prove that the relation $V$ in Figure 14 is also included in the simulation preorder. By Proposition 10 we have:

**Lemma 17.** For any $F$, $\rho$, and $C$,

$\psi[\phi[(\nu_{\iota_4}^{\alpha_{i_4}}{\delta_{i_4}}) (C \mid \text{TFS}([F], \emptyset, 0, [\rho]))]]$
\[\preceq(\nu_{\iota_4}^{\alpha_{i_4}}{\delta_{i_4}}) (C \mid \text{TFS}([F], \emptyset, 0, [\rho])) \]

So by Lemmas 16.1–2 and Corollary 8, $R$ is safe and fully abstract.

### 7 Designing secure distributed protocols

In the preceding sections, we present a thorough analysis of the problem of distributing access control. Let us now apply that analysis to a more general problem.

Suppose that we are required to design a distributed protocol that securely implements a specification. (The specification may be an arbitrary computation.) We can
solve this problem by partitioning the specification into smaller computations, running those computations in parallel, and securing the intermediate outputs of those computations so that they may be released and absorbed in any order. In particular, we can design $NS^{d+}$ by partitioning $IS^{d+}$ into access control and storage, running them in parallel, and securing the intermediate outputs of access control as capabilities. The same principles should guide any such design. For instance, by (R3) and (R4) intermediate outputs should not leak information prematurely; by (R5) and (R6) such outputs must be timestamped and the states on which they depend must not change between clock ticks; and by (M3) the specification must be generalized with time bounds.

**Computation as a graph** We describe a computation as a directed graph $G(V, E)$. The input nodes, collected by $V_i \subseteq V$, are the nodes of indegree 0. The output nodes, collected by $V_o \subseteq V$, are the nodes of outdegree 0. Further, we consider a set of state nodes $V_s \subseteq V$ such that $V_i \cap V_k = \emptyset$. As a technicality, any node that is in a cycle or has outdegree $> 1$ must be in $V_s$.

Nodes other than the input nodes run some code. Let $M$ contain all terms and $\sqsubseteq$ be a strict total order on $V$. We label each $v \in V \setminus (V_i \cup V_s)$ with a function $\lambda_v : M^{\text{in}(v)} \to M$, and each $v \in V_s$ with a function $\lambda_v : M^{\text{in}(v)} \times M \to M$. Further, each state node carries a shared clock, following the midnight-shift scheme.

A configuration $(\sigma, \tau)$ consists of a partial function $\sigma : V \to M$ such that $\text{dom}(\sigma) \supseteq V_s$, and a total function $\tau : V_s \to \mathbb{N}$. Intuitively, $\sigma$ assigns values at the state nodes and some other nodes, and $\tau$ assigns times at the state nodes. For any $v \in V \setminus V_i$, the function $\lambda_v$ outputs the value at $v$, taking as inputs the values at each incoming $u$, and the value at $v$ if $v$ is a state node; further, if such $u \notin V_s$, the value at $u$ is “consumed” on input. Formally, the operational semantics is given by a binary relation $\rightsquigarrow$ over configurations.

$$
\begin{align*}
v &\in V \setminus (V_i \cup V_s) & \forall k \in \text{in}(v). (u_k, v) &\in E & \sigma(u_k) &\in t_k \\
u_1 \sqsubseteq \ldots \sqsubseteq u_{\text{in}(v)} & & \sigma^- = \sigma|_{V_i \cup \{u_1, \ldots, u_{\text{in}(v)}\}} \\
\sigma, \tau &\rightsquigarrow (\sigma^-[v \mapsto \lambda_v(t_1, \ldots, t_{\text{in}(v)}), \tau])
\end{align*}
$$

$$
\begin{align*}
v &\in V_s & \tau(v) &\in \text{Clk} & \sigma(v) &\in t_k \\
u_1 \sqsubseteq \ldots \sqsubseteq u_{\text{in}(v)} & & \sigma^- = \sigma|_{V_i \cup \{u_1, \ldots, u_{\text{in}(v)}\}} \\
\sigma, \tau &\rightsquigarrow (\sigma^-[v \mapsto \lambda_v(t_1, \ldots, t_{\text{in}(v)}), \tau][v \mapsto \text{Clk} + 1])
\end{align*}
$$

As usual, we leave the context implicit; it can write values at $V_i$, read values at $V_o$, and read times at $V_s$.

For example, a graph that describes $IS^{d+}$ is:

$$
\begin{array}{ccccccc}
\bullet_1 & \rightarrow & \star_2 & \leftarrow & \star_4 & \rightarrow & \bullet_6 & \rightarrow & \star_7 & \rightarrow & \bullet_8 \\
\downarrow & & & & & & \uparrow & & & & & \\
\bullet_3 & \rightarrow & \star_4 & \rightarrow & \bullet_6 & \rightarrow & \star_7 & \rightarrow & \bullet_8
\end{array}
$$

Here $V_i = \{\bullet_1, \bullet_5\}$, $V_o = \{\bullet_3, \bullet_8\}$, $V_s = \{\star_2, \star_4, \star_7\}$, and $V = V_i \cup V_o \cup V_s \cup \{\bullet_6\}$. Intuitively, $\star_2$ carries accumulators, and $\bullet_1$ and $\bullet_3$ carry inputs and outputs for access
modifications; \( *_4 \) carries access policies, and \( \bullet_6 \) carries access decisions; \( *_7 \) carries stores, and \( \bullet_5 \) and \( \bullet_8 \) carry inputs and outputs for store operations. We define:

\[
\begin{align*}
\lambda_{\ast_2}(\langle k, \theta \rangle, F, (\_, \Xi)) &= \text{exec}(\text{perm}_{F,k,\theta}, \theta, \Xi) \\
\lambda_{\ast_3}(\langle N, \Xi \rangle) &= N \\
\lambda_{\ast_4}(\langle N, \Xi \rangle, \_ \rangle &= \Xi \\
\lambda_{\ast_6}(F, (k, op)) &= \langle op, \text{perm}_{F,k,op} \rangle \\
\lambda_{\ast_7}(\langle op, L \rangle, (\_, \rho)) &= \text{exec}(L, op, \rho) \\
\lambda_{\ast_8}(\langle N, \rho \rangle) &= N
\end{align*}
\]

**Distribution as a graph cut** Once described as a graph, a computation can be distributed along any cut of that graph. For instance, \( IS^{d+} \) can be distributed along the cut \( \{ \bullet_6, \ast_7 \} \) to obtain \( NS^{d+} \). We present this derivation formally in several steps.

**Step 1** For each \( v \in \mathcal{V} \), let \( S(v) \subseteq \mathcal{V}_s \) be the set of state nodes that have paths to \( v \), and \( I(v) \subseteq \mathcal{V}_i \) be the set of input nodes that have paths to \( v \) without passing through nodes in \( \mathcal{V}_s \). Then \( G(\mathcal{V}, E) \) can be written in a form where, loosely, the values at \( I(v) \) and the times at \( S(v) \) are explicit in \( \sigma(v) \) for each node \( v \). Formally, the *explication of \( G \) is the graph \( G(\mathcal{V}, E) \) where \( \mathcal{V}' = \mathcal{V} \cup \{ \hat{v} \mid v \in \mathcal{V}_i \} \cup \{ \hat{u} \mid u \in \mathcal{V}_o \} \) and \( \hat{E} = E \cup \{ (\hat{v}, v) \mid v \in \mathcal{V}_i \} \cup \{ (u, \hat{u}) \mid u \in \mathcal{V}_o \} \). We define:

\[
\begin{align*}
\lambda_v(t) &= \begin{cases} (t, t) & v \in \mathcal{V}_i \\
(\hat{t}, t) & v \in \mathcal{V}_o
\end{cases} \\
\lambda_v((I_1, t_1), \ldots, (I_{\text{in}(v)}, t_{\text{in}(v)})) &= \langle (I_1, \ldots, I_{\text{in}(v)}), t \rangle \\
\lambda_v((\_, t_1), \ldots, (\_, t_{\text{in}(v)}), (\_ \rangle, t) &= \langle Clk + 1, t' \rangle
\end{align*}
\]

This translation is sound and complete.

**Theorem 18.** \( \hat{G} \) is fully abstract with respect to \( G \).

For example, the explication of the graph for \( IS^{d+} \) is:

\[
\begin{array}{cccccccc}
\bullet_1 & \longrightarrow & *_2 & \longrightarrow & *_4 & \longrightarrow & \bullet_6 & \longrightarrow & *_7 & \longrightarrow & \bullet_8 \\
\downarrow & & \downarrow & & \uparrow & & \downarrow & & \uparrow & & \downarrow \\
\bullet_1 & & \bullet_3 & & \bullet_5 & & \bullet_8 \\
\downarrow & & \uparrow & & \downarrow & & \uparrow \\
\bullet_3 & & \bullet_5
\end{array}
\]

Here \( \sigma(\bullet_6) \) is of the form \( \langle k, op, Clk \rangle, \langle op, \text{perm}_{F,k,op} \rangle \) rather than \( \langle op, \text{perm}_{F,k,op} \rangle \); the “input” \( \sigma(\bullet_3) \) = \( \langle k, op \rangle \), the “time” \( \tau(*_4) \) = \( Clk \), and the “output” \( \langle op, \text{perm}_{F,k,op} \rangle \) of an access check are all explicit in \( \sigma(\bullet_6) \). A capability can be conveniently constructed from this form (see below).
Step 2 Next, let $\mathcal{E}_0$ be any cut. As a technicality, we assume that $\mathcal{E}_0 \cap (\mathcal{V}_t \cup \mathcal{V}_a) \times \mathcal{V}_i = \emptyset$. The distribution of $G$ along $\mathcal{E}_0$ is the graph $G^8(\mathcal{V}_8, \mathcal{E}_8)$, where $\mathcal{V}_8 = \mathcal{V} \cup \{ (v,\wang) \in \mathcal{E}_0 \} \cup \{ v^8 \mid (v,\wang) \in \mathcal{E}_0 \}$ and $\mathcal{E}_8 = (\mathcal{E} \setminus \mathcal{E}_0) \cup \{ (\pi, v^8) \mid (v,\wang) \in \mathcal{E}_0 \} \cup \{ (v^8, u) \mid (v,u) \in \mathcal{E}_0 \}$. Let $K_v$ and $E_v$ be secret keys shared by $v$ and $v^8$ for every $(v,\wang) \in \mathcal{E}_0$. We define:

\[
(v,\wang) \in \mathcal{E}_0 \quad \tilde{\lambda}_v(t_1, \ldots, t_{\text{in}(v)}) = \langle t, t' \rangle \quad m \text{ is fresh}
\]

\[
\lambda^6_v(t_1, \ldots, t_{\text{in}(v)}) = \text{mac}(\langle t, \{ m, t' \} \rangle_{E_v}, K_v)
\]

\[
(v,\wang) \in \mathcal{E}_0 \quad \tau(S(v)) \text{ is included in } t
\]

\[
\tilde{\lambda}^6_v(\langle t, \text{mac}(\langle t, \{ t' \} \rangle_{E_v}, K_v) \rangle) = \langle t, t' \rangle
\]

\[
v \in \mathcal{V} \setminus \mathcal{V}_i \quad (v,\wang) \notin \mathcal{E}_0
\]

\[
\lambda^8_v = \lambda_v
\]

Intuitively, for every $(v,\wang) \in \mathcal{E}_0$, $v^8$ carries the same values in $G^8$ as $v$ does in $G$; those values are encoded and released at $v$, absorbed at $\pi$, and decoded back at $v^8$. For example, the distribution of the graph for $IS^{d+}$ along the cut $\{ (\%_6, \%_7) \}$ is:

\[
\begin{array}{c}
\%_1 \rightarrow \%_2 \\
\uparrow \\
\%_1 \\
\%_3 \\
\downarrow \\
\%_3 \\
\%_6 \\
\%_5 \\
\uparrow \\
\%_5 \\
\uparrow \\
\%_5 \\
\uparrow \\
\%_5 \\
\end{array}
\]

This graph describes a variant of $NS^{d+}$. In particular, the node $\%_6$ now carries a capability of the form $\text{mac}(\langle \langle k, op, Clk \rangle, \{ m, \langle op, \text{perm}_{F,k,op} \} \rangle_{E_v}, K_{\%_6} \rangle)$, that secures the input, time, and output of an access check.

Step 3 Finally, $G$ is revised following $(A_5)$. The revision of $G$ along $\mathcal{E}_0$ is the graph $G^+(\mathcal{V}, \mathcal{E})$, where $\mathcal{V} = \mathcal{V} \cup \{ v^+ \mid (v,\wang) \in \mathcal{E}_0 \}$ and $\mathcal{E} = \mathcal{E} \setminus \{ (v^+, v) \mid (v,\wang) \in \mathcal{E}_0 \}$. We define:

\[
(v,u) \in \mathcal{E}_0 \quad \tau(S(v)) \leq T
\]

\[
\lambda^+(v(t_1, \ldots, t_{\text{in}(v)}, T)) = \lambda_v(t_1, \ldots, t_{\text{in}(v)})
\]

\[
v \in \mathcal{V} \setminus \mathcal{V}_i \quad (v,\wang) \notin \mathcal{E}_0
\]

\[
\lambda^+ = \lambda_v
\]

Intuitively, for every $(v,\wang) \in \mathcal{E}_0$, progress at $v$ requires that the times at $S(v)$ do not exceed the time bounds at $v^+$. For example, the revised form of the graph for $IS^{d+}$ is:

\[
\begin{array}{c}
\%_1 \rightarrow \%_2 \\
\uparrow \\
\%_3 \\
\%_6 \\
\%_5 \\
\%_5 \\
\%_8 \\
\%_8 \\
\%_8 \\
\end{array}
\]

Here $\%_6^+$ carries a time bound $T$, and $\lambda^+(\%_6^+, \langle k, op \rangle, T) = \lambda_{\%_6}(\%_6^+, \langle k, op \rangle)$ if $\tau(\%_5) \leq T$.

We prove the following correctness result.
Theorem 19. $G^8$ is fully abstract with respect to $G^\#$. By Theorem 19, the graph for $NS_d^+$ is fully abstract with respect to the revised graph for $IS_d^+$. Similarly, we can design $NS^*$ from $IS^*$. The induced subgraph of $IS_d^+$ without $\{\bullet_1, \star_2, \bullet_3, \star_4\}$ describes $IS^*$. We define $\lambda_\bullet_6(\langle k, op \rangle) = \langle op, \text{perm}_{F,k,op} \rangle$ for some static $F$. Distributing along the cut $\{(\bullet_6, \star_7)\}$, we obtain the induced subgraph of $NS_d^+$ without $\{\bullet_1, \bullet_1, \star_2, \bullet_3, \star_4\}$. This graph describes a variant of $NS^*$, with $\sigma(\bullet_6)$ of the form $\text{mac}(\langle \langle k, op \rangle, \{m, \langle op, \text{perm}_{F,k,op} \rangle \} \rangle, K_\bullet_6)$. (Here capabilities do not carry timestamps.) By Theorem 19 the graph for $NS^*$ is fully abstract with respect to a trivially revised graph for $IS^*$, where $\lambda_\bullet_6(\langle k, op \rangle) = \lambda_\bullet_6(\langle k, op \rangle)$.

8 Conclusion

We present a comprehensive analysis of the problem of implementing distributed access control with capabilities. In previous work, we show how to implement static access policies securely [10] and dynamic access policies safely [11]. In this paper, we explain those results in new light, revealing the several pitfalls that any such design must care about for correctness, while discovering interesting special cases that allow simpler implementations. Further, we present new insights on the difficulty of implementing dynamic access policies securely (a problem that has hitherto remained unsolved). We show that such an implementation is in fact possible if the specification is slightly generalized.

Moreover, our analysis turns out to be surprisingly general. Guided by the same basic principles, we show how to automatically derive secure distributed implementations of other stateful computations. This approach is reminiscent of secure program partitioning [22], and investigating its scope should be interesting future work.

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