Ensuring Consistency in Long Running Transactions

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ABSTRACT

Flow composition languages permit the construction of long-running transactions from collections of independent, atomic services. Due to environmental limitations, such transactions usually cannot be made to conform to standard ACID semantics. We propose *set consistency*, a powerful, yet intuitive, notion of consistency for long-running transactions. Set consistency specifications are written as predicates over the atomic actions of a process and require that the set of atomic actions executed along any system trace must satisfy the set consistency requirement. Set consistency generalizes *cancellation semantics*, a standard consistency requirement for long-running transactions, where failed processes are responsible for undoing any partially completed work, and can express strictly stronger requirements such as mutual exclusion or dependency. We believe that such predicates are both intuitive for developers and useful for representing real world constraints.

We formalize a core calculus for representing long-running transactions, providing sequential and parallel composition, as well as exception-handling and compensation actions, within the language. We show that the set consistency verification problem for processes in our core language against consistency requirements given as boolean predicates is co-NP complete and present an algorithm for verifying set consistency by reduction to propositional satisfiability. We have implemented this algorithm and demonstrate the value of our approach on three real-world case studies. In each case the consistency requirements can be verified within a second, demonstrating the practicality of our approach.

1. INTRODUCTION

In Web Services and Service Oriented Architectures, flow composition and orchestration languages [12, 1, 2] are used to combine disparate services into unified applications. Unlike monolithic applications built on a single database, such composite applications cannot take advantage of atomic transactions to ensure that changes are not partial or lost completely. Such applications generally recover using *compensation* — actions that are applied to undo other previously committed actions. Flow compensation languages may provide *compensation operators*, which associate compensation actions with completed actions and automatically run them in the event of a failure.

Existing work on using flow composition languages to implement transactions has focused on formalizing the semantics of compensation [6, 7, 8, 5, 4]. However, less effort has gone into defining a notion of correctness appropriate for long-running transactions. Standard, atomic transactions have a well-defined, application independent, criterion for correctness: the ACID (Atomicity, Consistency, Isolation, and Durability) properties [14]. However, ACID transactions are not suitable for flow composition. First, a transaction which composes web services is usually *long-running* — it takes a substantial amount of time to complete, because of asynchronous interactions across many systems, possibly requiring human intervention. Thus, the holding of locks and tight coordination of the participating systems is not feasible. Second, web services rarely expose distributed transaction interfaces, due to organizational boundaries, protocol limitations, and application limitations. Therefore, one must look for looser notions of correctness.

One such notion is *cancellation semantics* [8]: a long-running transaction should either complete successfully or undo all its observable results. The undo mechanism is a programmer-specified *compensation* action. For example, *sagas* [13] are two-level transactions with compensation actions for each forward action. Sagas achieve atomicity while relaxing the other requirements: either the entire transaction commits or any partial results are undone through running the compensation for each committed action.

Unfortunately, cancellation semantics is insufficient for many real-world scenarios. First, due to application requirements or the limitations of the actual systems, a process may contain non-compensatable actions and may need to continue execution even if certain actions fail. Thus, it is important to define a notion of correctness which accounts for "partial completion" of processes. Second, cancellation semantics cannot represent many application specific requirements, such as the mutual exclusion of two actions. With asynchronous interactions, parallel processing, and compensation, it is difficult to determine which actions actually completed.

We define an extended notion of correctness for long running transactions, called *set consistency*. Since the web services called by a flow composition are usually independent of each other, we consider only whether services are called, not their respective execution orderings. Like Sagas, we assume individual service calls are atomic and thus treat failure the same as if a service was never called. Rather than defining correctness with respect to a process trace, we define it with respect to an *execution* — the set of actions which completed successfully during a run of a process. A *set consistency specification* is a set of executions corresponding to correct runs of a process. We show how such a specification can be compactly represented using propositional logic.

Set consistency specifications can represent cancellation semantics as well as stronger restrictions on such processes. In addition, they can specify behaviors which relax the self-cancellation requirement. If the web services called by a flow composition are independent, set consistency can capture many requirements currently modeled using temporal logic. We believe that set consistency predicates are more intuitive than temporal logic specifications for developers.

We formalize the notion of set consistency using a core calculus for process composition. Our core language composes atomic actions using sequential and parallel composition, choice, compensation, and exception handling constructs, and is an abstraction of common flow composition languages such as BPEL. Then, the *set consistency verification problem* checks that all feasible executions of a process are contained in the set of executions defined by the set consistency specification. We show that this problem is co-NP complete and present an algorithm for verifying set consistency by constructing a predicate representing all feasible process executions. This reduces the verification problem to propositional validity which can be checked using an off-the-shelf SAT solver.

In contrast to temporal logic model checking of processes, which is undecidable [11], set consistency is a powerful, yet tractable class of properties that can already capture many interesting notions of correctness. We have implemented the algorithm for set consistency verification and applied it to three case studies, including a real-world CRM system which contained a consistency bug and was the inspiration of our work (see Section 2). The verification problems resulting from our case studies can each be discharged within a second, showing that the technique is feasible for real-world applications.

This is a longer version of our conference paper. It includes outlines for the proofs and extended examples.

2. EXAMPLE

We now informally describe *set consistency* using an example inspired by a bug actually seen in a production system. One of the authors of this paper previously worked for an enterprise applications vendor. Once, when visiting a customer (a large bank), the author reviewed a set of the customer's business processes, which integrated a mainframe-based financial application with a CRM (Customer Relationship Management) application. When an account was created or changed in the financial application, a message containing the account was sent to the CRM system via a transactional queue. The CRM system took the message from the queue and updated its version of the account accordingly. When examining the business process run by the CRM system, the author found a consistency bug, which inspired the work in this paper.

Upon taking a message from the queue, the CRM system executed a business process which transformed the message to its internal account representation, performed other bookkeeping, and saved the account. If the account was saved successfully, the CRM system should commit its transaction with the queue, to avoid duplicate messages. If the account save failed, the CRM system should abort its transaction with the queue, to avoid losing the update entirely. Upon the abort of a message "take", the message was automatically put back on the queue, unless a retry threshold was exceeded. In this case, the message went to a special "dead-letter" queue, to be investigated by an administrator. We can model the interactions of the CRM system with the queue as the following process:

$$AcctRecv = \textbf{TakeMsg}; (body \rhd (\textbf{Abort}; \textsf{throw})); \textbf{Commit}$$

Here, **bold fonts** represent atomic actions (the implementation of *body* is not shown). The ";" operator composes two actions or processes sequentially: it runs the first process, and if it is successful, runs the second. The "$\rhd$" operator catches exceptions: it runs the first process, and, only if it fails, runs the second. throw is a special built-in action which throws an exception. The **TakeMsg** process takes a message off the queue. The subprocess *body* handles each account. If *body* fails, then the message transaction is aborted (by calling **Abort**), putting the message back on the queue. Otherwise, **Commit** commits the message transaction, permanently removing it from the queue. Ignoring all the implementation details of *body*, we want to ensure that, if the action **SaveAcct** is called within *body*, then **Commit** is called, and if **SaveAcct** is not called, then **Abort** is called.

Deep within the *body* subprocess, the author found the equivalent of the following code:

$$\textbf{SaveAcct} \rhd \textbf{LogErr}$$

where **SaveAcct** performs the actual account write. Someone had added an exception handler which, in the event of an error in saving, logged a debug message. Unfortunately, they left this code in their production system. The exception handler "swallows" the error status — **Commit** will always be called, even if the save fails, violating our correctness requirement. If, due to a data validation error or a transient system problem, the CRM system rejects an account, bank tellers will not be able to find the customer!

Traces. Executing a process gives rise to a trace. A *trace* is the sequence of successfully executed actions invoked by the process, along with the final status of the run (either $\checkmark$ or $\times$). For example, if the **SaveAcct** action fails, we get the following trace:

$$T = \textbf{TakeMsg}, \textbf{Preprocess}, \textbf{LogErr}, \textbf{Commit}\langle\checkmark\rangle$$

The **Preprocess** action represents the preprocessing which occurs in *body* before **SaveAcct** is called. The error from **SaveAcct** is not propagated to the outer process, and thus the queue transaction is committed. Note that the failed invocation of **SaveAcct** does not appear between **Preprocess** and **LogErr**. We leave out failed invocations as they have no permanent, observable effect on the system.

Set Consistency. Our first observation is that the bug in *AcctRecv* can be caught if we know that there exists a feasible trace in which **Commit** is executed and **SaveAcct** is either not called or is invoked, but failed. That is, we can abstract away the relative order of individual actions, and only consider the *set* of actions that executed successfully. Accordingly, we define an *execution* as the set of actions which appear in a trace. The execution E associated with the above trace is:

$$E = \{\textbf{TakeMsg}, \textbf{Preprocess}, \textbf{LogErr}, \textbf{Commit}\}$$

A *set consistency specification* defines a set of "good" executions. For example, correct executions of our account process either:

1. include both **SaveAcct** and **Commit**, or
2. include **Abort**, but not **SaveAcct**.

We write set consistency specifications in a predicate notation, where the literal a means that action a is included in the execution, and $\neg a$ means that action a is not included in the execution and literals are combined using the boolean

∧ (and), ∨ (or), and ¬ (not) operators. A *set consistency predicate* is interpreted over the actions which appear in the process. It represents a set consistency specification which contains exactly those executions that satisfy the predicate. To represent the above two conditions, we can define the specification predicate φ_{q1} as follows:

$$\varphi_{q1} = (\textbf{SaveAcct} \wedge \textbf{Commit}) \vee (\neg\textbf{SaveAcct} \wedge \textbf{Abort})$$

Other consistency predicates can be specified for our process. For example, we might want to ensure that **Abort** and **Commit** are never called in the same run:

$$\varphi_{q2} = (\textbf{Abort} \wedge \neg\textbf{Commit}) \vee (\neg\textbf{Abort} \wedge \textbf{Commit})$$

We can check both requirements simultaneously by taking the conjunction of the two predicates: $\varphi_{q3} = \varphi_{q1} \wedge \varphi_{q2}$.

Verification. The *set consistency verification problem* takes as input a process P and a consistency predicate φ and asks if all feasible executions of P satisfy φ. If so, we say that the process *satisfies* the specification. Clearly, the process *AcctRecv* does not satisfy the specification φ_{q1} — a counterexample is the execution E above, which clearly does not satisfy φ_{q1}. However, the process does meet the specification φ_{q2}.

In Section 4, we show that one can check a specification by constructing a predicate ϕ_p that represents all the feasible executions for the process P. Then, P satisfies the specification φ if ϕ_p implies φ, which is a propositional satisfiability check. We have built a verifier, described in section 6, that verifies set consistency by this method. If we run *AcctRecv* through our verifier, using specification φ_{q3}, it returns the execution E above as a counterexample.

Note that we can fix the problem by either removing the exception handler in *body* or re-throwing the exception after calling **LogErr**. With either of these fixes, the process passes our verifier.

Other Features. In practice, business process flows contain, in addition to sequencing and exception handling, parallel composition (where processes execute in parallel) and *compensations* (processes that get executed to undo the effect of atomic actions should a subsequent process fail). Our process language (and verifier) allows us to write both parallel composition and compensation actions.

As an example, consider an alternative version of *body* which saves the account in two steps. First, it writes a header using the action **SaveHdr** and then writes contact information using the action **AddContact**. Both of these actions can fail. If **SaveHdr** succeeds, but **AddContact** fails, we undo the account change through the compensating action **DelHdr**, which never fails. In our process language, we specify a compensation action using the "÷" operator. Our new process is written as:

$$AcctRecv_2 = \textbf{TakeMsg}; (body_2 \triangleright (\textbf{Abort}; \textsf{throw})); \textbf{Commit}$$
$$body_2 = (\textbf{SaveHdr} \div \textbf{DelHdr}); \textbf{AddContact}$$

A set consistency requirement for $AcctRecv_2$ states that if both the account header and contact actions are successful, then **Commit** should be run. Otherwise, **Abort** should be run. That is,

$$(\textbf{SaveHdr} \wedge \textbf{AddContact} \wedge \textbf{Commit}) \vee$$
$$(\neg\textbf{SaveHdr} \wedge \neg\textbf{AddContact} \wedge \textbf{Abort})$$

The negated action ¬**SaveHdr** captures the scenarios where either **SaveHdr** is not run, or it is run but subsequently compensated by **DelHdr**. We capture the effect of compensations using a programmer-defined *normalization*

Atomic Action	A	::=	skip	throw	built-ins
				$A_i \in \mathcal{A}$	defined actions
Process	P	::=	A		atomic actions
				P; P	sequence
				P ∥ P	parallel
				P □ P	choice
				P ÷ P	compensation
				P ▷ P	exception handler

Figure 1: Syntax of process calculus.

set $\mathcal{C} = \{(\textbf{SaveHdr}, \textbf{DelHdr})\}$ that encodes that the effect of running **SaveHdr** can be undone by running its compensation **DelHdr**. Given the predicate and the normalization set, our verifier automatically expands the predicate to

$$(\textbf{SaveHdr} \wedge \neg\textbf{DelHdr} \wedge \textbf{AddContact} \wedge \textbf{Commit}) \vee$$
$$((\neg\textbf{SaveHdr} \vee (\textbf{SaveHdr} \wedge \textbf{DelHdr})) \wedge \neg\textbf{AddContact} \wedge \textbf{Abort})$$

This expanded specification predicate is then used by our algorithm, which shows that process $AcctRecv_2$ satisfies this set consistency specification.

3. PROCESS CALCULUS

We use a simple process calculus to describe long-running transactions. The operators in our calculus correspond to similar concrete operations in web service orchestration languages like BPEL.

Syntax. Figure 1 defines the syntax for our language. Processes are constructed from *atomic actions*, using a set of composition operations. Atomic actions are indivisible operations which either succeed completely or fail and undo all state changes. There are two built-in actions: skip, which always succeeds and does nothing, and throw, which throws an exception. We use $\mathcal{A}$ for the set of atomic actions defined by the environment and Σ for the set of all atomic actions: $\Sigma \equiv \mathcal{A} \cup \{\textsf{throw}, \textsf{skip}\}$.

A process is either an atomic action or a composition of atomic actions using one of five composition operators. The sequence operator ";" runs the first process followed by the second. If the first fails, the second is not run. The parallel operator "∥" runs two processes in parallel. The choice operator "□" non-deterministically selects one of two processes and runs it. The compensation operator "÷" runs the left process. If it completes successfully, the right process is installed as a compensation to run if the parent process terminates with an error. The exception handler "▷" runs the left process. If that process terminates with an error, the error is ignored and the right process is run. If the left process terminates successfully, the right process is ignored. Our core language does not include iteration operators —we will add iteration to our language in section 5.

In our examples, we also use *named subprocesses* — subprocesses which are defined once and then appear as atomic actions in the overall process — as syntactic sugar to enhance readability. Named subprocesses are not true functions — they are simply inlined into their parent. Thus, recursive calls are not permitted.

We define $|P|$, the size of process P, by induction: the size $|A|$ of an atomic action in Σ is 1, and the size $|P_1 \otimes P_2|$ for any composition operation $\otimes$ applied to P_1 and P_2 is $|P_1| + |P_2| + 1$.

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Process form	Π
$\Gamma \vdash A : \{\checkmark, \times\}$	$\{(A\langle\checkmark\rangle, \mathsf{skip}), (\langle\times\rangle, \mathsf{skip})\}$
$\Gamma \vdash A : \{\checkmark\}$	$\{(A\langle\checkmark\rangle, \mathsf{skip})\}$
$\Gamma \vdash A : \{\times\}$	$\{(\langle\times\rangle, \mathsf{skip})\}$
P; Q	$\{(pq\langle s\rangle, Q'; P') \mid (p\langle\checkmark\rangle, P') \in \Pi(P), (q\langle s\rangle, Q') \in \Pi(Q)\} \ \cup \ \{(p\langle\times\rangle, P') \mid (p\langle\times\rangle, P') \in \Pi(P)\}$
P ∥ Q	$\{(r\langle s \& t\rangle, P' \parallel Q') \mid r \in p \bowtie q, (p\langle s\rangle, P') \in \Pi(P), (q\langle t\rangle, Q') \in \Pi(Q)\}$ $\cup \ \{(p\langle\times\rangle, P') \mid (p\langle\times\rangle, P') \in \Pi(P)\} \ \cup \ \{(q\langle\times\rangle, Q') \mid (q\langle\times\rangle, Q') \in \Pi(Q)\}$
P □ Q	$\{(p\langle s\rangle, P') \mid (p\langle s\rangle, P') \in \Pi(P)\} \ \cup \ \{(q\langle t\rangle, Q') \mid (q\langle t\rangle, Q') \in \Pi(Q)\}$
P ÷ Q	$\{(p\langle\checkmark\rangle, Q) \mid (p\langle\checkmark\rangle, P') \in \Pi(P)\} \ \cup \ \{(p\langle\times\rangle, P') \mid (p\langle\times\rangle, P') \in \Pi(P)\}$
P ▷ Q	$\{(p\langle\checkmark\rangle, P') \mid (p\langle\checkmark\rangle, P') \in \Pi(P)\} \ \cup \ \{(pp'q\langle t\rangle, Q') \mid pp'\langle\times\rangle \in [\![P]\!], (q\langle t\rangle, Q') \in \Pi(Q)\}$

Figure 2: Trace semantics.

Trace Semantics. We now define a trace-based semantics for processes. A *run* is a (possibly empty) sequence of atomic actions from Σ. A *trace* is a run followed by either $\langle\checkmark\rangle$ or $\langle\times\rangle$, representing successful and failed executions, respectively. For example, if action A_1 is run successfully and then action A_2 fails, the corresponding trace would be $A_1\langle\times\rangle$. In the following, we let the variable A range over atomic actions; the variables P, P', Q, Q', and R range over processes; the variables p, q, and r range over runs, and the variables s and t range over $\checkmark$ and $\times$.

We define an operator "&" which combines two process status symbols ($\checkmark$, $\times$). Given $s\&t$, if both s and t are $\checkmark$, then $s\&t = \checkmark$. Otherwise, $s\&t = \times$. For the parallel composition rules, we introduce an operator $\bowtie: \mathcal{R} \times \mathcal{R} \to 2^{\mathcal{R}}$, where, given two runs p and q, $p \bowtie q$ produces the set of all interleavings of p and q.

We use the symbol Γ to represent an *action type environment*, which maps each action to a set of possible results $2^{\{\checkmark, \times\}} = \{\{\checkmark\}, \{\times\}, \{\checkmark, \times\}\}$ from running the associated action. This allows us to distinguish actions which may fail from actions which never fail. The result set $\{\times\}$ is for the special action throw, which unconditionally throws an error.

For each process form P, we define using mutual induction two semantic functions $\Pi(P)$ and $[\![P]\!]$. The rules for $\Pi : \mathcal{P} \to 2^{(\mathcal{T}, \mathcal{P})}$, found in figure 2, define a set of pairs. Each pair $(p\langle s\rangle, Q) \in \Pi(P)$ consists of a trace $p\langle s\rangle$, representing a possible execution of the process P, and a process Q, representing a *compensation process* associated with the trace that will get run on a subsequent failure.

The function $[\![P]\!] : \mathcal{P} \to 2^{\mathcal{T}}$, maps a process P to a set of feasible traces of P. To compute the actual traces possible for a top-level process, this function must consider the successful and failed traces independently. Compensation processes are dropped from successful traces. For a failed trace $p\langle\times\rangle$, one computes all possible compensation traces for the associated compensation process P' and appends these to p. $[\![P]\!]$ is defined as follows:

$$[\![P]\!] = \{(p\langle\checkmark\rangle | (p\langle\checkmark\rangle, P') \in \Pi(P)\} \ \cup$$
$$\{pp'\langle\times\rangle | (p\langle\times\rangle, P') \in \Pi(P), p'\langle s\rangle \in [\![P']\!]\}$$

We now describe the rules in Figure 2 in more detail.

Atomic Actions. For atomic actions, we enumerate the possible results for each action. Individual atomic actions use skip as a compensation process, since no compensation is provided by default. Compensation for an atomic action must be explictly defined using the $\div$ operator.

Sequential Composition. For a sequential composition $P; Q$, we consider two cases. If P succeeds, we have a successful trace $p\langle\checkmark\rangle$ for P, after which we concatenate a trace q from Q. The status for the overall trace is then $\langle s\rangle$, the status from the trace q. The compensation process (if Q fails) invokes the compensation for Q first, followed by the compensation for P. On the other hand, if P fails, we have a failed trace $p\langle\times\rangle$ for P. The process Q is not run at all.

Parallel Composition. For parallel composition, we first consider the case where both sub-processes run. If both are successful, the entire process is successful. If one fails, the process throws an error. We simulate the parallel semantics by generating a possible trace for all interleavings of the two subprocesses. The compensation for the two sub-processes is also run in parallel. Note that, if a sub-process fails, the other sub-process runs to completion, unless it also encounters an error. However, if the second sub-process has not started yet, and the first fails, an implementation can avoid running the second at all. This is handled by the last two sets in the union.

Choice Composition. The traces for $P \square Q$ are simply the union of the traces for P and Q.

Compensation. The compensation operator $P \div Q$ runs P and then sets up process Q as compensation. If P is successful, Q overrides any previous compensation for P — e.g. P' if $\Pi(P) = (p\langle\checkmark\rangle, P')$. If P fails, then the original compensation process returned by $\Pi(P)$ is instead returned. In this case, the process Q is never run.

Exception Handler. The rule for the exception handling operator has two scenarios. Given $P \triangleright Q$, if P is successful, Q is ignored. If P fails, the compensation for P is run and then process Q is run. To express this, we use $[\![P]\!]$ to obtain a full trace for P, including compensation. This makes the $[\![\cdot]\!]$ and Π functions mutually recursive.

Example 1. To demonstrate these rules, we compute the

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set of feasible traces for the *SimpleOrder* process defined as:

$$SimpleOrder = Billing; \textbf{ProcessOrder}$$
$$Billing = \textbf{Charge} \div \textbf{Credit}$$

This process first calls the *Billing* sub-process. This sub-process invokes the **Charge** action to bills the customer's action. If this fails, the process terminates. If the charge succeeds, **Credit** is registered as a compensation action. Next, **ProcessOrder** is run to handle the actual order. If **ProcessOrder** is successful, the process terminates successfully. Otherwise, the **Charge** compensation is run and the process terminates with an error.

We assume that the **Charge** and **ProcessOrder** atomic actions both can fail, but the **Credit** compensation never fails. Thus, we obtain the following definition for Γ:

$$\langle \textbf{Charge} \mapsto \{\checkmark \times\}, \textbf{Credit} \mapsto \{\checkmark\}, \textbf{ProcessOrder} \mapsto \{\checkmark, \times\}\rangle$$

If we apply the rules for atomic actions to each action in this process, we get the following values for Π:

$$\Pi(\textbf{Charge}) = \{(\textbf{Charge}\langle\checkmark\rangle, \mathsf{skip}), (\langle\times\rangle, \mathsf{skip})\}$$
$$\Pi(\textbf{Credit}) = \{(\textbf{Credit}\langle\checkmark\rangle, \mathsf{skip})\}$$
$$\Pi(\textbf{ProcessOrder}) = \{(\textbf{ProcessOrder}\langle\checkmark\rangle, \mathsf{skip}), (\langle\times\rangle, \mathsf{skip})\}$$

To compute the feasible traces for the *Billing* subprocess, we apply the rule for compensation, with $P = \textbf{Charge}$ and $Q = \textbf{Credit}$, obtaining the following:

$$\Pi(Billing) = \{(\textbf{Charge}\langle\checkmark\rangle, \textbf{Credit}), (\langle\times\rangle, \mathsf{skip})\}$$

To compute the feasible traces for *SimpleOrder* we use the sequential composition rule, using $P = Billing$ and $Q = \textbf{ProcessOrder}$:

$$\Pi(SimpleOrder) = \{(\textbf{Charge ProcessOrder}\langle\checkmark\rangle, \mathsf{skip}; \textbf{Credit}),$$
$$(\textbf{Charge}\langle\times\rangle, \textbf{Credit}), (\langle\times\rangle, \mathsf{skip})\}$$

Finally, we compute the feasible traces $[\![SimpleOrder]\!]$ by dropping compensation for successful traces and computing the compensation traces for failed traces:

$$\{\textbf{Charge ProcessOrder}\langle\checkmark\rangle, \textbf{Charge Credit}\langle\times\rangle, \mathsf{skip}\langle\times\rangle\}$$

Trace composition. We end this section by looking at how the composition of processes affects the composition of traces. Given a process R and trace $T_r = r_f r_c \langle s\rangle$, where $T_r \in [\![R]\!]$, $(r_f \langle s\rangle, R') \in \Pi(R)$, and $r_c \langle s'\rangle \in [\![R']\!]$, we call the trace $r_f \langle s\rangle$ the *forward sub-trace* of T_r and $r_c \langle s'\rangle$ the *compensation sub-trace* of T_r. We write $r_f r_c \langle s\rangle$ for a trace r with forward sub-trace r_f and compensation sub-trace r_c and status $\langle s\rangle$. Note that the overall status of the trace is always equal to the status of the forward sub-trace.

Let $R = P \otimes Q$, where $\otimes$ is one of the composition operators. For each feasible trace T_r of R, either P is not called (e.g., if $\otimes = \Box$ and Q is chosen), or there is a trace T_p of P such that the forward subtrace of P occurs within the forward subtrace of T_r and the compensation subtrace of T_p appears in the compensation subtrace of T_r.

THEOREM 1. *Let* $R = P \otimes Q$, $p_f p_c \langle t\rangle \in [\![P]\!]$, *and* $r_f r_c \langle s\rangle \in [\![R]\!]$ *such that* $r_f = r_{f_0} p_f r_{f_1}$ *(the run* p_f *appears within* r_f*), and* $r_c = r_{c_0} p_c r_{c_1}$ *(the run* p_c *appears within* r_c*). The forward sub-trace* $r_{f_1}\langle s\rangle$ *of* R *following the call to* P *depends only on* $\langle t\rangle$ *(the forward status of* P*), not on the individual actions in* p_f*. Likewise, the compensation sub-trace* $r_{c_1}\langle s'\rangle$ *of* R *following the call to* P' *depends only on the compensation status of* P'*, not on the individual compensation actions in* p_c*.*

PROOF OUTLINE: The proof is by induction on the derivation of $[\![R]\!]$, using a case analysis on the syntactic forms of each subprocess. For each composition operator, inspection of the corresponding rules of Figure 2 show that, in each case, there is no dependency on the individual actions called by the subprocesses and the status of the parent process depends only upon the status of the subprocesses. $\square$

4. SET CONSISTENCY

We can now formalize our notion of consistency with respect to our process calculus.

Executions. For a trace $p\langle s\rangle \in [\![P]\!]$, we define the *execution* for $p\langle s\rangle$ as the set $\pi_p \subseteq \Sigma$ of atomic actions that appear in p, that is, the execution $e(p) = \{a \in \Sigma \mid \exists p_1, p_2. p\langle s\rangle \equiv p_1 a p_2 \langle s\rangle\}$. For a process P, let $\mathsf{execs}(P) \subseteq 2^\Sigma$ represent the set of all executions of P, defined by $\mathsf{execs}(P) = \{e(p) \mid p\langle s\rangle \in [\![P]\!]\}$.

Example 2. If we drop calls to skip, the set of feasible executions $\mathsf{execs}(SimpleOrder)$ for *SimpleOrder* is

$$\{\{\textbf{Charge}, \textbf{ProcessOrder}\}, \{\textbf{Charge}, \textbf{Credit}\}, \emptyset\}$$

Set Consistency Specifications. A *set consistency specification* $\mathcal{S} \subseteq 2^\Sigma$ is a set of action sets representing the permissible executions for a given process. A process P is *set consistent* with respect to a set consistency specification $\mathcal{S}$ if all executions of P fall within the set consistency specification: $\mathsf{execs}(P) \subseteq \mathcal{S}$.

Set consistency specifications are semantic objects. We use a boolean predicate-based syntax for describing sets of executions. Given a set of atomic actions Σ, a *set consistency predicate* is an expression built from combining atomic predicates Σ with the logical operators $\neg$ (not), $\wedge$ (and), and $\vee$ (or). The size $|\varphi|$ of predicate φ is the defined by induction: $|a| = 1$ for a literal, and $|\varphi_1 \otimes \varphi_2| = |\varphi_1| + |\varphi_2| + 1$ for a logical operator $\otimes$. To evaluate a predicate φ on an execution $e \in 2^\Sigma$, we assign the value $\mathtt{true}$ to all atomic actions that occur in e and $\mathtt{false}$ to all the atomic actions that do not occur in e. If the predicate φ evaluates to $\mathtt{true}$ with these assignments, we say the execution e *satisfies* the specification φ, and write $e \models \varphi$. The set consistency specification $\mathcal{S}$ defined by the set consistency predicate φ is the set of satisfying assignments of φ:

$$\mathsf{spec}(\varphi) = \{e \in 2^\Sigma \mid e \models \varphi\}$$

Normalization. When defining set consistency specifications, we wish to treat the compensated execution of an action as equivalent to an execution where the action (and its compensation) never occurs. If A° is a compensation action for A, then the term $\neg A$ in the specification predicate should yield two executions in the final, expanded specification: one with neither A nor A°, and one with both A and A°. We call a specification that has been adjusted in this manner a *normalized* specification. Normalization is performed with respect to a programmer-specified *normalization set* $\mathcal{C} \subseteq \{(a, a^\circ) | a, a^\circ \in \Sigma\}$ of atomic action pairs, where the second action in each pair is the compensation action for the first action. Given the consistency specification predicate φ and a normalization set $\mathcal{C}$, we apply the function $\mathsf{spec\_norm}(\varphi, \mathcal{C})$. This function uses DeMorgan's laws and the normalization set to convert the predicate to a form where (1) the negation operator only appears in front

of literals, and (2) given a pair (a, a') from $\mathcal{C}$, each occurrence of $\neg a$ is replaced with $(\neg a \wedge \neg a') \vee (a \wedge a')$ and each occurrence of a is replaced with $a \wedge \neg a'$.

For successful executions of an action, the normalized specification asserts that the compensation action was not run. For failed executions, spec_norm changes the specification to treat equivalently the following three scenarios:

1. Actions a and a' are never run.
2. Action a is executed but fails, undoing any partial state changes. Action a' is never run.
3. Action a is executed successfully, but compensation a' is later run to undo the effects of a.

Example 3. We consider a specification for the *SimpleOrder* process. We assume the action typing Γ_{so} of example 1 (in which the **Credit** action never fails) and the normalization set $\mathcal{C}_{so} = \{(\textbf{Charge}, \textbf{Credit})\}$.

We wish to have cancellation semantics, where either both **Charge** and **ProcessOrder** complete successfully, or, in a failed execution, any completed actions are undone. Given the normalization set, we do not need to distinguish between failure cases. Thus, we can write a set consistency predicate φ_{so} for *SimpleOrder* as:

$$(\textbf{Charge} \wedge \textbf{ProcessOrder}) \vee (\neg\textbf{Charge} \wedge \neg\textbf{ProcessOrder})$$

We now expand φ_{so} to get spec_norm$(\varphi_{so}, \mathcal{C}_{so})$. It is already in the form we need for substitution (negation only of literals). We replace all occurrences of $\neg\textbf{Charge}$ with $(\neg\textbf{Charge} \wedge \neg\textbf{Credit}) \vee (\textbf{Credit} \wedge \textbf{Charge})$ and all occurrences of **Charge** with $(\textbf{Charge} \wedge \neg\textbf{Credit})$:

$$(\textbf{Charge} \wedge \neg\textbf{Credit} \wedge \textbf{ProcessOrder}) \vee$$

$$((\neg\textbf{Charge} \wedge \neg\textbf{Credit}) \vee (\textbf{Charge} \wedge \textbf{Credit})) \wedge \neg\textbf{ProcessOrder}$$

Set consistency verification. For a process P, a normalization set $\mathcal{C}$, and a set consistency predicate φ, the set consistency *verification problem* is to check if all the executions of P w.r.t. $\mathcal{C}$ satisfy φ, that is, if execs$(P) \subseteq$ spec(spec_norm$(\varphi, \mathcal{C})$).

THEOREM 2. *The set consistency verification problem is co-NP complete.*

PROOF OUTLINE: Verification is in co-NP, as finding a counterexample is in NP. To find a counterexample, we nondeterministically enumerate feasible executions of a process using the trace semantics. At each possible decision point (e.g. success or failure of actions and the choice operator), we guess an outcome. We track only the set of actions which were called successfully, and, upon termination, check whether this execution is a satisfying assignment of the specification predicate. An execution is polynomial in the size of the process, since it includes each action at most once. Checking whether the execution satisfies the specification predicate φ is polynomial with respect to $|\varphi|$.

To show that verification is co-NP hard, we reduce checking for tautology to set consistency verification. To determine whether a predicate ϕ, consisting of literals from the set Σ_ϕ, is a tautology, we first construct a process P whose feasible executions are the powerset of Σ_ϕ. One such process is $A_1 \parallel A_2... \parallel A_n$ where $A_1, ...A_n \in \Sigma_\phi$ and $\forall i \, . \, \Gamma \vdash A_i : \{\checkmark, \times\}$. Now, we interpret ϕ as a set consistency predicate: If process P satisfies the specification ϕ, then ϕ is a tautology. $\square$

Predicate-based verification. We now give an algorithm for verifying set consistency. We define a syntax-directed analysis which takes a process as input and constructs a predicate ϕ whose satisfying assignments precisely represent the feasible execution set. The predicate, ϕ is composed from atomic predicates Σ using logical operators $\neg$, $\wedge$, $\vee$, $\rightarrow$, and $\leftrightarrow$. For the moment, we assume that a given action A is referenced only once in a process. Later, we will extend our approach to remove this restriction. A predicate ϕ over atomic predicates in Σ defines a set of executions $\mathcal{E}(\phi) = \{e \in 2^\Sigma \mid e \models \phi\}$.

Execution Predicate Construction. We create the predicate ϕ by recursively applying the rules of Figure 3, based on the form of each sub-process. These rules define seven mutually-recursive functions: $\phi_{\checkmark 0}$, $\phi_{\checkmark\checkmark}$, $\phi_{\checkmark\times}$, $\phi_{\times 0}$, $\phi_{\times\checkmark}$, $\phi_{\times\times}$, and ϕ_{00}, all with the signature $\mathcal{P} \times \mathcal{G} \rightarrow \Phi$, where $\mathcal{P}$ is the set of all processes, $\mathcal{G}$ is the set of all action type environments, and Φ the set of all predicates. The two subscripts of each function's name represent the forward and compensation results of running the process, respectively, where $\checkmark$ is a successful execution, $\times$ is a failed execution, and 0 means that the process was not run. For example, $\phi_{\checkmark\times}(P, \Gamma)$ returns a predicate representing all executions of process P where the forward process completes successfully and the compensation process (whose execution must be initiated by the failure of a containing process) fails, given a type environment Γ. As a shorthand, we use terms of the form $P_{\checkmark\times}$ to represent the function $\phi_{\checkmark\times}(P, \Gamma)$. We also leave out the conjunction symbol ("$\wedge$") when it is obvious from the context (e.g. $P_{\checkmark 0}Q_{\checkmark 0}$ for $P_{\checkmark 0} \wedge Q_{\checkmark 0}$).

Given these functions, we compute the predicate ϕ, representing the possible executions of a process, using the function pred : $\mathcal{P} \times \mathcal{G} \rightarrow \Phi$, defined as:

$$\text{pred}(P, \Gamma) \equiv \phi_{\checkmark 0}(P, \Gamma) \vee \phi_{\times\checkmark}(P, \Gamma) \vee \phi_{\times\times}(P, \Gamma)$$

Example 4. We illustrate the predicate generation algorithm by constructing a predicate for the *SimpleOrder* process. We start from the execution predicate definition:

$$\text{pred}(SimpleOrder, \Gamma_{so}) = \phi_{\checkmark 0} \vee \phi_{\times\checkmark} \vee \phi_{\times\times}$$

and iteratively apply the appropriate sub-predicates from Figure 3. We show the steps of the computation in Figure 4. We obtain step (2) from step (1) in Figure 4 by substituting each term with the sub-predicates from the sequence rule, since the top-level process is a sequence. Next, from step (2) to step (3), we use the atomic action rules to simplify the predicates for **ProcessOrder**. Notice that the fourth conjunction in (3) can be dropped since it conjoins false. Now, we expand the predicates for *Billing* using rules for the compensation operator to get step (4). Finally, we expand the **Charge** and **Credit** predicates again using rules for the atomic actions to get (5). Simplifying, we get (6), the execution predicate for *SimpleOrder*.

Note that the three conjunctions in the final predicate correspond to the three possible executions of *SimpleOrder*:

1. The actions **Charge** and **ProcessOrder** are successful, and the compensation **Credit** is not run.
2. The action **Charge** completes successfully, but **ProcessOrder** fails. Then, **Credit** is run to compensate for **Charge**.
3. The action **Charge** fails, and the actions **ProcessOrder** and **Credit** are never run.

Memoization. As usual, one can memoize each computation by giving names to sub-predicates. While generat-

Predicate	$\Gamma \vdash A : \{\checkmark\}$	$\Gamma \vdash A : \{\checkmark, \times\}$	$\Gamma \vdash A : \{\times\}$
$\phi_{\checkmark 0}$	A	A	`false`
$\phi_{\checkmark\checkmark}$	A	A	`false`
$\phi_{\checkmark\times}$	`false`	`false`	`false`
$\phi_{\times 0}$	`false`	$\neg A$	$\neg A$
$\phi_{\times\checkmark}$	`false`	$\neg A$	$\neg A$
$\phi_{\times\times}$	`false`	`false`	`false`
ϕ_{00}	$\neg A$	$\neg A$	$\neg A$

Predicate	**P;Q**	**P∥Q**	**P□Q**
$\phi_{\checkmark 0}$	$P_{\checkmark 0}Q_{\checkmark 0}$	$P_{\checkmark 0}Q_{\checkmark 0}$	$P_{00}Q_{\checkmark 0}\vee\ P_{\checkmark 0}Q_{00}$
$\phi_{\checkmark\checkmark}$	$P_{\checkmark\checkmark}Q_{\checkmark\checkmark}$	$P_{\checkmark\checkmark}Q_{\checkmark\checkmark}$	$P_{00}Q_{\checkmark\checkmark}\vee\ P_{\checkmark\checkmark}Q_{00}$
$\phi_{\checkmark\times}$	$P_{\checkmark 0}Q_{\checkmark\times}\vee\ P_{\checkmark\times}Q_{\checkmark 0}$	$P_{\checkmark 0}Q_{\checkmark\times}\vee\ P_{\checkmark\checkmark}Q_{\checkmark\times}\vee\ P_{\checkmark\times}Q_{\checkmark 0}\vee$ $P_{\checkmark\times}Q_{\checkmark\checkmark}\vee\ P_{\checkmark\times}Q_{\checkmark\times}$	$P_{00}Q_{\checkmark\times}\vee\ P_{\checkmark\times}Q_{00}$
$\phi_{\times 0}$	$P_{\checkmark 0}Q_{\times 0}\vee\ P_{\times 0}Q_{00}$	$P_{00}Q_{\times 0}\vee\ P_{\checkmark 0}Q_{\times 0}\vee\ P_{\times 0}Q_{00}\vee$ $P_{\times 0}Q_{\checkmark 0}\vee\ P_{\times 0}Q_{\times 0}$	$P_{00}Q_{\times 0}\vee\ P_{\times 0}Q_{00}$
$\phi_{\times\checkmark}$	$P_{\checkmark\checkmark}Q_{\times\checkmark}\vee\ P_{\times\checkmark}Q_{00}$	$P_{00}Q_{\times\checkmark}\vee\ P_{\checkmark\checkmark}Q_{\times\checkmark}\vee\ P_{\times\checkmark}Q_{00}\vee$ $P_{\times\checkmark}Q_{\checkmark\checkmark}\vee\ P_{\times\checkmark}Q_{\times\checkmark}$	$P_{00}Q_{\times\checkmark}\vee\ P_{\times\checkmark}Q_{00}$
$\phi_{\times\times}$	$P_{\checkmark 0}Q_{\times\times}\vee\ P_{\checkmark\times}Q_{\times\checkmark}\vee\ P_{\times\times}Q_{00}$	$P_{00}Q_{\times\times}\vee\ P_{\checkmark 0}Q_{\times\times}\vee\ P_{\checkmark\checkmark}Q_{\times\times}\vee$ $P_{\checkmark\times}Q_{\times 0}\vee\ P_{\checkmark\times}Q_{\times\checkmark}\vee\ P_{\checkmark\times}Q_{\times\times}\vee$ $P_{\times 0}Q_{\checkmark\times}\vee\ P_{\times 0}Q_{\times\times}\vee\ P_{\times\times}Q_{\checkmark\times}\vee$ $P_{\times\checkmark}Q_{\times\times}\vee\ P_{\times\times}Q_{00}\vee\ P_{\times\times}Q_{\checkmark 0}\vee$ $P_{\times\times}Q_{\checkmark\checkmark}\vee\ P_{\times\times}Q_{\checkmark\times}\vee\ P_{\times\times}Q_{\times 0}\vee$ $P_{\times\times}Q_{\times\checkmark}\vee\ P_{\times\times}Q_{\times\times}$	$P_{00}Q_{\times\times}\vee\ P_{\times\times}Q_{00}$
ϕ_{00}	$P_{00}Q_{00}$	$P_{00}Q_{00}$	$P_{00}Q_{00}$

Predicate	**P÷Q**	**P▷Q**
$\phi_{\checkmark 0}$	$P_{\checkmark 0}Q_{00}$	$P_{\checkmark 0}Q_{00}\vee\ P_{\times\checkmark}Q_{\checkmark 0}$
$\phi_{\checkmark\checkmark}$	$P_{\checkmark 0}Q_{\checkmark 0}$	$P_{\checkmark\checkmark}Q_{00}\vee\ P_{\times\checkmark}Q_{\checkmark\checkmark}$
$\phi_{\checkmark\times}$	$P_{\checkmark 0}Q_{\times\checkmark}\vee\ P_{\checkmark 0}Q_{\times\times}$	$P_{\checkmark\times}Q_{00}\vee\ P_{\times\checkmark}Q_{\checkmark\times}$
$\phi_{\times 0}$	$P_{\times 0}Q_{00}$	$P_{\times\checkmark}Q_{\times 0}$
$\phi_{\times\checkmark}$	$P_{\times\checkmark}Q_{00}$	$P_{\times\checkmark}Q_{\times\checkmark}$
$\phi_{\times\times}$	$P_{\times\times}Q_{00}$	$P_{\times\checkmark}Q_{\times\times}\vee\ P_{\times\times}Q_{00}$
ϕ_{00}	$P_{00}Q_{00}$	$P_{00}Q_{00}$

Figure 3: Inference rules for computing ϕ

ing the execution predicate, we name each non-atomic sub-predicate, using the names as literals instead of expanding the sub-predicates of ϕ. The resulting execution predicate is then conjoined with a definition $(n \leftrightarrow \phi_n)$ for each name n and its definition ϕ_n.

Using memoization, the execution predicate of our previous example becomes:

$$(SimpleOrder_{\checkmark 0} \vee SimpleOrder_{\times\checkmark})$$
$$\wedge\ (SimpleOrder_{\checkmark 0} \leftrightarrow Billing_{\checkmark 0} \wedge \textbf{ProcessOrder})$$
$$\wedge\ (Billing_{\checkmark 0} \leftrightarrow \textbf{Charge} \wedge \neg\textbf{Credit})$$
$$\wedge\ (SimpleOrder_{\times\checkmark} \leftrightarrow$$
$$(Billing_{\checkmark\checkmark} \wedge \neg\textbf{ProcessOrder} \vee$$
$$Billing_{\times\checkmark} \wedge \neg\textbf{ProcessOrder}))$$
$$\wedge\ (Billing_{\checkmark\checkmark} \leftrightarrow \textbf{Charge} \wedge \textbf{Credit})$$
$$\wedge\ (Billing_{\times\checkmark} \leftrightarrow \neg\textbf{Charge} \wedge \neg\textbf{Credit})$$

Although the memoized execution predicate for this particular example is larger than the non-memoized predicate, in general, the worst-case size of the memoized predicate is polynomial in the size of a process, whereas the size of a non-memoized predicate can be exponential.

THEOREM 3. *For any process P and action type environment Γ,*
1. *The execution set obtained from P's execution predicate is equal to the set of all P executions: $\mathcal{E}(\textsf{pred}(P, \Gamma)) = \textsf{execs}(P, \Gamma)$.*
2. *$|\textsf{pred}(P, \Gamma)|$ is polynomial in $|P|$.*

PROOF OUTLINE:

Execution predicate soundness. We prove the first assertion by induction on the derivation of $\textsf{pred}(P)$, first considering direct generation of the execution predicate, without memoization. The base cases are each of the atomic action rules. For each rule, we show that the satisfying assignments to the predicates correspond to the feasible executions. The inductive cases are the compositional operator rules. We can simplify the proof for these cases using theorem 1, which shows that, when composing subprocesses, one can compose the traces, only considering the status (success or failure) of the subtraces. For each operator, we show that, assuming that the component predicates (e.g. $P_{\checkmark\checkmark}$, etc.) are correct, then the predicate constucted for the operator has satisfying assignments corresponding to the executions obtained from the associated trace semantics rule.

Memoization. To prove that our memoization algorithm is correct, we show a more general property for an arbitary predicate ϕ containing a sub-predicate φ. An assignment $\mathbb{A}$ of values to the variables of ϕ is a satisfying assignment for ϕ, if and only if it is satisfying assignment for $(\phi[C/\varphi]) \wedge (C \leftrightarrow \varphi)$, where C is a literal not appearing in ϕ. To see this, we use several boolean algebra identities. We convert $\phi[C/\varphi]$ to disjunctive normal form, yielding ϕ_{dnf}.

We assume without proving here that replacing each occurrance of C with φ in ϕ_{dnf} would yield a predicate equivalent to ϕ. Given this, we only need to show that conjoining ϕ_{dnf} with $C \leftrightarrow \varphi$ is equivalent to performing this substitution. We conjoin each term t_i of ϕ_{dnf} with $C \wedge \varphi \vee \neg C \wedge \neg\varphi$, creating terms t_i^C and $t_i^{\neg C}$. If C and not $\neg C$ appear in t_i,

$$\mathsf{pred}(SimpleOrder, \Gamma_{so}) \;=\; \phi_{\checkmark 0} \;\vee\; \phi_{\times\checkmark} \;\vee \phi_{\times\times} \tag{1}$$

$$\begin{aligned}
=\; & Billing_{\checkmark 0}\mathbf{ProcessOrder}_{\checkmark 0} \;\vee\; Billing_{\checkmark\checkmark}\mathbf{ProcessOrder}_{\times\checkmark} \;\vee\; \\
& Billing_{\times\checkmark}\mathbf{ProcessOrder}_{00} \;\vee\; Billing_{\checkmark 0}\mathbf{ProcessOrder}_{\times\times} \;\vee\; \\
& Billing_{\checkmark\times}\mathbf{ProcessOrder}_{\times\checkmark} \;\vee\; Billing_{\times\times}\mathbf{ProcessOrder}_{00}
\end{aligned} \tag{2}$$

$$\begin{aligned}
=\; & Billing_{\checkmark 0} \wedge \mathbf{ProcessOrder} \;\vee\; Billing_{\checkmark\checkmark} \wedge \neg\mathbf{ProcessOrder} \;\vee\; \\
& Billing_{\times\checkmark} \wedge \neg\mathbf{ProcessOrder} \;\vee\; Billing_{\checkmark 0} \wedge \mathtt{false} \;\vee\; \\
& Billing_{\checkmark\times} \wedge \neg\mathbf{ProcessOrder} \;\vee\; Billing_{\times\times} \wedge \neg\mathbf{ProcessOrder}
\end{aligned} \tag{3}$$

$$\begin{aligned}
=\; & \mathbf{Charge}_{\checkmark 0} \wedge \mathbf{Credit}_{00} \wedge \mathbf{ProcessOrder} \;\vee\; \mathbf{Charge}_{\checkmark 0} \wedge \mathbf{Credit}_{\checkmark 0} \wedge \neg\mathbf{ProcessOrder} \;\vee\; \\
& \mathbf{Charge}_{\times\checkmark} \wedge \mathbf{Credit}_{00} \wedge \neg\mathbf{ProcessOrder} \;\vee\; \\
& (\mathbf{Charge}_{\checkmark 0}\mathbf{Credit}_{\times\checkmark} \vee \mathbf{Charge}_{\checkmark 0}\mathbf{Credit}_{\times\times}) \wedge \neg\mathbf{ProcessOrder} \;\vee\; \\
& \mathbf{Charge}_{\times\times} \wedge \mathbf{Credit}_{00} \wedge \neg\mathbf{ProcessOrder}
\end{aligned} \tag{4}$$

$$\begin{aligned}
=\; & \mathbf{Charge} \wedge \neg\mathbf{Credit} \wedge \mathbf{ProcessOrder} \;\vee\; \mathbf{Charge} \wedge \mathbf{Credit} \wedge \neg\mathbf{ProcessOrder} \;\vee\; \\
& \neg\mathbf{Charge} \wedge \neg\mathbf{Credit} \wedge \neg\mathbf{ProcessOrder} \;\vee\; (\mathbf{Charge} \wedge \mathtt{false} \vee \mathbf{Charge} \wedge \mathtt{false}) \wedge \neg\mathbf{ProcessOrder} \;\vee\; \\
& \mathtt{false} \wedge \neg\mathbf{Credit} \wedge \neg\mathbf{ProcessOrder}
\end{aligned} \tag{5}$$

$$\begin{aligned}
=\; & \mathbf{Charge} \wedge \neg\mathbf{Credit} \wedge \mathbf{ProcessOrder} \;\vee\; \mathbf{Charge} \wedge \mathbf{Credit} \wedge \neg\mathbf{ProcessOrder} \;\vee\; \\
& \neg\mathbf{Charge} \wedge \neg\mathbf{Credit} \wedge \neg\mathbf{ProcessOrder}
\end{aligned} \tag{6}$$

Figure 4: Computing $\mathsf{pred}(SimpleOrder, \Gamma_{so})$

then $t_i^C = t_i \wedge \varphi$ and $t_i^{\neg C}$ is unsatisfiable. When considering satisfiability, we can ignore the term C in t_i^C, since it is a free variable in ϕ_{dnf}. The case where t_i contains $\neg C$ but not C is similar. If t_i contains C and $\neg C$, then t_i and $t_i[\varphi/C]$ are unsatisfiable. Thus, $\phi_{dnf} \wedge (C \leftrightarrow \varphi)$ is equivalent to $\phi_{dnf}[\varphi/C]$.

Execution predicate size. To show that the execution predicate's size is polynomial with respect to the original process, we use induction on the structure of processes. The base causes are atomic actions which all have predicates of one term. For the inductive case, we consider an abitrary subprocess $P \otimes Q$ where the predicates $\phi_{\checkmark 0}(P), ...\phi_{00}(P)$ are all polynomial in the size of P and the predicates $\phi_{\checkmark 0}(Q), ...\phi_{00}(Q)$ are all polynomial in the size of Q. Using memoization, we need only reference each of these predicates once. The predicates $\phi_{\checkmark 0}(P \otimes Q), ...\phi_{00}(P \otimes Q)$ are thus independent in size from the sizes of P and Q. The largest possible size is, in fact, 67 for $\phi_{\times\times}(P \parallel Q)$.

Checking Consistency. We check a specification by checking if the execution predicate $\mathsf{pred}(P, \Gamma)$ implies the normalized specification predicate $\mathsf{spec\_norm}(\varphi, \mathcal{C})$. If the implication is valid, then all executions satisfy the specification and the solution to the consistency verification problem is "yes." Otherwise, there is some execution that does not satisfy the consistency specification. Therefore, to check a process for consistency, we can build the execution and normalized specification predicates and check the implication by a Boolean satisfiability query.

THEOREM 4. *For any process P and specification predicate φ, with action type environment Γ and normalization set $\mathcal{C}$,*
1. $\mathsf{execs}(P, \Gamma) \subseteq \mathsf{spec}(\mathsf{spec\_norm}(\varphi, \mathcal{C}))$ *iff* $\mathsf{pred}(P, \Gamma) \rightarrow \mathsf{spec\_norm}(\varphi, \mathcal{C})$ *is valid.*
2. *The consistency verification problem can be solved in time exponential in some polynomial function of $|P|$, $|\Gamma|$, $|\varphi|$, and $|\mathcal{C}|$.*

PROOF OUTLINE: This follows from the previous theorems. If the execution predicate is a sound and complete representation of a process's executions, then the process satisifies a specification only if the execution predicate implies the specification.

The verification predicate can be generated in polynomial time with respect to the algorithm's inputs: only one pass must be made through the process and the size of the generated predicate is polynomial with respect to $|P|$, $|\Gamma|$, $|\varphi|$, and $|\mathcal{C}|$. From theorem 3 we know that the execution predicate is polynomial in size with respect to P. $|\Gamma|$ and $|\mathcal{C}|$ are not larger than the number of unique actions in P and $|\phi|$ is independent from Γ. Finally, $\mathsf{spec\_norm}$ causes at most a polynomial expansion of the specification predicate. $\square$

Example 5. Given the previously computed values of $\mathsf{pred}(SimpleOrder, \Gamma_{so})$ and $\mathsf{spec\_norm}(\varphi_{so}, \mathcal{C}_{so})$, the verification problem for process $SimpleOrder$ may be reduced to checking the validity of the following predicate:

$$\begin{aligned}
& (\mathbf{Charge} \wedge \neg\mathbf{Credit} \wedge \mathbf{ProcessOrder} \;\vee\; \\
& \;\;\mathbf{Charge} \wedge \mathbf{Credit} \wedge \neg\mathbf{ProcessOrder} \;\vee\; \\
& \;\;\neg\mathbf{Charge} \wedge \neg\mathbf{Credit} \wedge \neg\mathbf{ProcessOrder}) \;\rightarrow\; \\
& ((\mathbf{Charge} \wedge \mathbf{ProcessOrder}) \;\vee\; \\
& \;\;((\neg\mathbf{Charge} \wedge \neg\mathbf{Credit}) \vee (\mathbf{Charge} \wedge \mathbf{Credit})) \\
& \;\;\;\;\wedge \neg\mathbf{ProcessOrder})
\end{aligned}$$

Multiple Calls to an Action. So far, we have assumed that each action in $\mathcal{A}$ is called at most once in a given process. If an action may be called multiple times by a process, we do not distinguish the individual calls. Given an action A, which appears more than once in the text of the process P, the specification predicate A is $\mathtt{true}$ if A is called at least once, and $\mathtt{false}$ if A is never called. This follows directly from our definition of a process execution, which is a *set* of called actions, rather than a *bag*. These semantics are useful for situations where an action may be called in

one of several situations, and we wish to verify that, given some common condition, the action is called. For example, in $OrderProcess$ of section 6.1, the action `MarkFailed` must be called if the order fails, either due to a failed credit check or to a failed fulfillment subprocess.

If an action is called more than once, $\mathsf{pred}(P, \Gamma)$ may produce an unsatisfiable predicate. For example, $\phi_{\checkmark 0}(A \square A) = (A \wedge \neg A) \vee (\neg A \wedge A)$, which is clearly unsatisfiable. We solve this by applying the function $\mathsf{trans\_calls} : \mathcal{P} \rightarrow \mathcal{P} \times \Phi$, which takes as input a process P and returns a translated process P' along with a predicate ϕ_{tc}. Given a set of actions $\{A^1, ... A^n\} \subseteq \mathcal{A}$, which occur more than once in P, each occurrence of these actions the translated process P' is given a unique integer subscript. For example, given the process $P = (A; B) \square (A; B)$, $\mathsf{trans\_calls}(P)$ will return the process $P' = (A_1; B_1) \square (A_2; B_2)$. The predicate $\phi_{tc} = \phi_{tc}^1 \wedge ... \phi_{tc}^n$, where ϕ_{tc}^i uses the boolean $\leftrightarrow$ operator to associate the atomic predicate A^i with the disjunction of the predicates for each subscript. For example, ϕ_{tc} for $P = (A; B) \square (A; B)$ will be $(A \leftrightarrow (A_1 \vee A_2)) \wedge (B \leftrightarrow (B_1 \vee B_2))$. We now re-define our predicate function pred, to combine $\mathsf{trans\_calls}$ with $\mathsf{pred_{mem}}$, our original predicate generation function:

$$\mathsf{pred}(P, \Gamma) = \texttt{let } (P', \phi_{tc}) = \mathsf{trans\_calls}(P) \texttt{ in}$$

$$\mathsf{pred_{mem}}(P', \Gamma) \wedge \phi_{tc}$$

LEMMA 1. *For any process P, which may contain multiple calls to the same action, and atomic action typing Γ, the execution set obtained from applying pred to P and Γ is equal to the set of all P executions: $\mathcal{E}(\mathsf{pred}(P, \Gamma)) = \mathsf{execs}(P, \Gamma)$.*

PROOF OUTLINE: We show this through boolean algebra identities. Assume that process P contains an action A which is called n times. Thus, $\mathsf{trans\_calls}(P) = (P', A \leftrightarrow (A_1 \vee A_2... \vee A_n))$ and $\phi_{p'} = \mathsf{pred_{mem}}(P')$, where P' has actions $A_1, A_2, ... A_n$ substituted for each occurrence of A. We convert $\phi_{p'}$ to disjunctive normal form, yielding $\phi_{p'}^{dnf}$. We state, without proving here, that each term of $\phi_{p'}^{dnf}$ will contain all actions of P', either directly or in negated form. We conjoin each term t_i of $\phi_{p'}^{dnf}$ with $A \leftrightarrow (A_1 \vee ... A_n)$. This yields the terms $t_i^{A_1} = t_i \wedge A \wedge A_1$, $t_i^{A_2} = t_i \wedge A \wedge A_2$, ... $t_i^{A_n} = t_i \wedge A \wedge A_n$, and $t_i^{\neg A} = t_i \wedge \neg A \wedge \neg A_1 ... \wedge \neg A_n$. The terms $t_i^{A_j}$ are satisifiable if t_i satisfiable, A is assigned `true`, and t_i contains A_j as a conjunct (and thus not $\neg A_j$). The term $t_i^{\neg A}$ is satisfiable if t_i is satisfiable, A is assigned `false`, and t_i contains $\neg A_1, ... \neg A_n$ as conjuncts. If we view the terms t_i as possible executions for P, this matches our intuition: if any of the subscripted actions A_j are included in the execution, then we include A. $\square$

Named subprocesses. As mentioned in section 3, named subprocesses are simply expanded at each callpoint. This is done before running $\mathsf{trans\_calls}$, since expansion may introduce actions which are called more than once.

5. ITERATION

We now extend our core process language to support iteration by introducing two new process composition operators, "$**$" and "$*|$", with the following syntax:

```
Process  P  ::=  ...
            |  **P    sequential iteration
            |  *|P    parallel iteration
```

The sequential iteration operator runs a process one or more times, one copy at a time. The parallel iteration operator runs one or more copies of a process in parallel. For both operations, the number of copies is not known a priori — this is determined at runtime.

We define the trace semantics of these operators by taking the fixpoint of the following two equations:

$$\Pi(* * P) = \Pi(P) \ \cup \ \Pi(P; (* * P))$$
$$\Pi(*|P) = \Pi(P) \ \cup \ \Pi(P \parallel (*|P))$$

The sequential iteration operator generates a sequence of traces from $\Pi(P)$, where all but the last trace must correspond to successful executions. As with pairwise sequential composition, the compensation process for sequential iteration runs the compensation processes for each action sequentially, in reverse order from the forward execution. Parallel set iteration interleaves traces of P arbitrarily. If any of these traces fails, the parent has a failed status. As with parallel composition, processes may not be interrupted in the event of a failure, but may be skipped if they have not started. Compensation is run in parallel as well.

We wish to generalize set consistency predicates to processes with iteration. We encounter two problems with iteration. First, the set of potential traces for a given process now becomes infinite. Second, consider two atomic actions A and B, which occur within an iterated subprocess. The specification $A \wedge B$ will be true for all traces where both A and B are called at least once, even if A and B are never called in the same iteration. This is usually too weak. In an executable process, an iterated subprocess would likely be parameterized by some value (dropped when abstracting to our core language), which changes each iteration. Thus, one is more likely interested in checking if, whenever A is called, B is also called within the same iteration. We will see an example of such a process in section 6.1.

Quantified specification predicates. We now generalize the semantics of specification predicates to permit specifying that all iterations of an subprocess obey a constraint. As a running example, we use the following process:

$$P = (* * (A; B)); (*|(C \square D))$$

First, we augment process traces to uniquely identify each call to an action within iterative subprocesses. We number the iterations of each iterative subprocess, starting at 1. We add this iteration number as a subscript to the actions called within each iteration. For example, given an run of process P which executes the $A; B$ iteration twice and runs two copies of C and one copy of D for the parallel iteration, we obtain the trace: $A_1 \ B_1 \ A_2 \ B_2 \ C_1 \ C_2 \ D_3$. For parallel iteration, since subprocesses may be interleaved, the mapping of subprocess executions to iterations is arbitrary. If an iterative subprocess is nested within another iterative subprocess, we subscript each inner subprocess's actions with a sequence of iteration numbers, starting with the outermost subprocess. We extend the trace semantics function $\llbracket \cdot \rrbracket$ to produce traces of this form for iterative subprocesses.

The execution for an iterative trace is simply the set of actions called, propagating the iteration subscripts. Thus, the actions of each iteration are distinguished. For example, execution for the above trace of process P would be: $\{A_1, A_2, B_1, B_2, C_1, C_2, D_3\}$. We extend the execution function $\mathsf{execs} : \mathcal{P} \rightarrow 2^{\mathcal{A}}$ to produce executions of this form.

Next, we extend the specification semantics function $\mathsf{spec} : \Phi \times \mathcal{P} \to \mathcal{E}$ to permit specification predicates over iterative executions. As discussed above, we wish to consider whether a property holds over all iterations. Therefore, we forbid the use of specific iteration subscripts in the specification predicate. Instead, we assign a unique index variable to each iterative subprocess and universally quantify the specification predicate over all index variables. For example, to specify that, for each sequential iteration in process P, the execution of B implies the execution of A, we write $\forall_i(\neg B_i \vee A_i)$. We write $\varphi^\forall$ to represent a quantified specification predicate. Predicates outside any iteration are not subscripted. Thus, given the process $P' = (* * (X; Y)); Z$, the predicate $\forall_i(\neg Z \vee (X_i \wedge Y_i))$, means "in all executions of the process where Z is called, X and Y are called together for all iterations."

We can now define the semantics of an execution predicate, given a universally quantified specification predicate $\varphi^\forall$ and a process P:

$$\mathsf{spec}^\forall(\varphi^\forall, P) = \{e \in \mathsf{execs}(P) \mid e \models \varphi^\forall\}$$

Verification. Clearly, if a quantified specification predicate is satisfied by all possible iterations of a process, then any single iteration will satisfy the predicate as well. More importantly, the reverse is also true. To show this, we first define a function $\mathsf{erase} : \mathcal{P} \to \mathcal{P}$, which removes the iteration operators from a process, replacing them with the underlying subprocess:

$$\mathsf{erase}(P) = \begin{cases} \mathsf{erase}(P') & \text{if } P = * * (P') \\ \mathsf{erase}(P') & \text{if } P = *|(P') \\ \mathsf{erase}(Q) \otimes \mathsf{erase}(R) & \text{if } P = Q \otimes R \\ A & \text{if } P = A \end{cases}$$

Given a specification predicate $\varphi^\forall$, we write φ^ϵ for the predicate obtained by removing the universal quantifier and any iteration subscripts on literals.

THEOREM 5. *Given a process P, a normalization set $\mathcal{C}$, and quantified specification predicate $\varphi^\forall$, iff $\mathsf{erase}(P)$ is correct with respect to φ^ϵ, then P is correct with respect to $\varphi^\forall$.*

PROOF OUTLINE: Assume that process P contains an iterative subprocess $Q = * * (R)$ or $Q = *|(R)$.

The function $\mathsf{select\_iters} : 2^{\mathcal{A}} \to \mathcal{E}$ takes an execution E of P and returns the set of executions $E^1, E^2, ... E^n$ where each E^i substitutes the execution E_{q_i} of Q iteration i for the execution E_q of all Q iterations. We can lift $\mathsf{select\_iters}$ to sets of executions by applying it to each execution in the input set and accumulating all resulting executions in a single set.

We may evaluate a feasible execution set $\mathsf{execs}(P)$ of process P against the quantified execution predicate $\varphi^\forall$ by evaluating each execution in $\mathsf{select\_iters}(\mathsf{execs}(P))$ against φ^ϵ. If all these executions satisfy φ^ϵ, then P satisfies $\varphi^\forall$.

LEMMA 2.

$$\forall_{e \in \mathsf{select\_iters}(\mathsf{execs}(P))} \cdot (e \models \varphi^\epsilon) \to$$
$$\forall_{e' \in \mathsf{execs}(P)} \cdot (e' \models \varphi^\forall)$$

From the trace definitions for iteration, we can see that the possible traces for an single iteration of Q are the same

Process	Spec	Proc size	Spec size	Pred size	Time (ms)
AccountReceive	φ_{q1}	13	8	124	120
OrderProcess	φ_{o1}	26	21	247	124
OrderProcess	φ_{o2}	26	20	236	130
BrokenOrder	φ_{o2}	22	20	198	157
Travel	φ_{t1}	13	24	181	119
Travel	φ_{t2}	13	39	210	127

Figure 5: Experimental Results. "Spec" is the consistency specification, "Spec size" is the size of the specification, "Pred size" is the size of the execution predicate.

as if the subprocess R is run standalone. In other words, iterations are "stateless." Thus, the possible executions and forward/compensation status pairs for any one iteration of Q are the same as those for R. From theorem 1, we know that the trace of the process calling Q depends only on the status of Q, not the individual actions. Thus, the set of feasible executions we get when replacing Q with R are the same as when we select each iteration from the executions of P. More formally:

LEMMA 3. $\mathsf{execs}(\mathsf{erase}(P)) = \mathsf{select\_iters}(\mathsf{execs}(P))$.

Theorem 5 follows from lemmas 2 and 3. □

Thus, we can check an iterative process by simply erasing the iterative operators, and checking against the associated unquantified specification.

6. EXPERIENCES

To demonstrate the viability of our approach, we have implemented a verifier for our process language, based on the predicate-generation algorithm of section 4. We then modeled the example of section 2 and two additional case studies in our language and verified them using our tool. The source code for our verifier and examples may be found at `http://cs.ucla.edu/~fischer/code/trans`.

We implemented the predicate generation algorithm in Objective Caml. To determine the validity of verification predicates, we use the Simplify theorem prover [9]. The predicate generator and Simplify are separate executables —the predicate is passed to Simplify via an operating system pipe. While this is inefficient, and Simplify's SAT capabilities are far behind modern SAT solvers, the performance of our verifier has not been a problem in our case studies.

Figure 5 shows the results of running our verifier on the example of section 2 and the two case studies below. The run times are for a 1.6Ghz Pentium M laptop with 512 MB of memory, running Windows XP. The runs all complete in less than 160 milliseconds. The verification predicates are approximately 10 times larger than the original processes. Since process languages are intended for composing lower-level services, the size of real-world processes are usually not more than an order of magnitude larger than our examples, well within the capabilities of SAT solvers.

6.1 Case Study: Order Process

[8] presents a simple order fulfillment transaction which has cancellation semantics, and thus compensation is sufficient for expressing the required error handling. Figure 6 shows a more complex order fulfillment transaction, inspired

$$OrderProcess = \textbf{SaveOrder}; CreditCheck;$$
$$\textbf{SplitOrder}; FulfillOrder; CompleteOrder$$
$$CreditCheck = (\textbf{ReserveCredit} \rhd (\textbf{Failed}; \mathsf{throw}))$$
$$\div OrderFailed$$
$$OrderFailed = \textbf{RestoreCredit}; \textbf{Failed}$$
$$FulfillOrder = * | (ProcessPO_i)$$
$$ProcessPO_i = (\textbf{FulfillPO}_i \rhd (\textbf{MarkPOFailed}_i; \mathsf{throw}))$$
$$\div \textbf{CancelPO_i}$$
$$CompleteOrder = \textbf{BillCustomer}; \textbf{Complete}$$

Figure 6: Order management process

by the example application in [18], which does not have cancellation semantics.

The process *OrderProcess* receives a customer's order and makes a reservation against the customer's credit (**ReserveCredit**). If the customer does not have enough credit, the order is marked *failed* (**Failed**) and processing stopped. Otherwise, if the credit check was successful, the subprocess *OrderFailed* is installed as compensation for the credit check. Then, the order is broken down into sub-orders by supplier, and these sub-orders are submitted to their respective suppliers in parallel (subprocess *FulfillOrder*). Upon completion of the sub-orders, if all are successful, the subprocess *CompleteOrder* finalizes the credit transaction with a call to **BillCustomer** and marks the order as *complete* (**Complete**).

Failure semantics. There are two types of errors that may occur in the order transaction. If the *ReserveCredit* call fails (e.g. due to insufficient credit), the order is marked as failed and execution is terminated before submitting any purchase orders. Alternatively, one or more purchase orders may be rejected by the associated suppliers. If any orders fail, the credit reservation is undone and the order marked as failed. Note that neither error scenario causes the entire transaction to be undone. This is consistent with real world business applications, where many transactions have some notion of partial success and records of even failed transactions are retained.

We assume the normalization set

$$\{(\textbf{ReserveCredit}, \textbf{RestoreCredit}), (\textbf{FulfillPO}, \textbf{CancelPO})\}$$

and that the actions **SaveOrder**, **RestoreCredit**, **CancelPO**, **Failed**, and **Complete** never fail.

The order should always be saved. If the order process is successful, the customer should be billed, all the purchase orders fulfilled, and the order marked complete. If the order process fails, the order should be marked as failed, the customer should not be billed, and no purchase orders fulfilled. These requirements are written as the following set consistency predicate φ_{o1}:

SaveOrder $\wedge$

$((\textbf{BillCustomer} \wedge \textbf{FulfillPO} \wedge \textbf{Complete} \wedge \neg\textbf{Failed}) \vee$

$(\neg\textbf{BillCustomer} \wedge \neg\textbf{FulfillPO} \wedge \neg\textbf{Complete} \wedge \textbf{Failed}))$

When checked with our verifier, *OrderProcess* does indeed satisfy this specification.

Next, we consider an alternative, orthogonal specification. Assume that the **ReserveCredit**, **RestoreCredit**,

and **BillCustomer** actions all belong to an external credit service. We wish to ensure that our process always leaves the service in a consistent state: if **ReserveCredit** succeeds, then either **RestoreCredit** or **BillCustomer** (but not both) must eventually be called. Also, if **ReserveCredit** fails, neither should be called. We model these requirements with the predicate φ_{o2}:

$(\neg\textbf{ReserveCredit} \rightarrow (\neg\textbf{RestoreCredit} \wedge \neg\textbf{BillCustomer})) \wedge$
$(\textbf{ReserveCredit} \rightarrow (\textbf{RestoreCredit} \oplus \textbf{BillCustomer}))$

where $\rightarrow$ and $\oplus$ are syntactic sugar for logical implication and logical exclusive-or, respectively. Since we are referencing **RestoreCredit** directly in our specification, we remove the pair (**ReserveCredit**, **RestoreCredit**) from our compensation set. Our verifier finds that *OrderProcess* satisfies this specification.

Finally, we consider the process *BrokenOrder*, a variation of *OrderProcess* where the *OrderFailed* compensation is left out of the *CreditCheck* subprocess:

$$CreditCheck = \textbf{ReserveCredit} \rhd (\textbf{Failed}; \mathsf{throw})$$

When checking this process, our verifier finds that the φ_{o2} specification is not satisfied and returns the following counter-example execution:

$$\{\textbf{SaveOrder}, \textbf{ReserveCredit}, \textbf{SplitOrder}, \textbf{Failed}\}$$

This execution corresponds to a trace where the process runs successfully until it reaches **FulfillPO**, which fails. The exception handling for **FulfillPO** runs **Failed**, but **RestoreCredit** is never run to undo the effects of **ReserveCredit**.

6.2 Case Study: Travel Agency

Many real world applications involve *mixed transactions*. A mixed transaction combines both compensatable and non-compensatable subtransactions [10]. Frequently, these processes involve a *pivot action* [17], which cannot be compensated or retried. To obtain cancellation semantics, actions committing before the pivot must support compensation (backward recovery) and actions committing after the pivot must either never fail or support retry (forward recovery).

Set consistency specifications can capture these requirements and our verifier can check these properties. To illustrate this, we use a travel agency example from [15]. We model it in our core language as follows:

$$Travel = ((\textbf{BookFlight} \div \textbf{CancelFlight};$$
$$(\textbf{RentCar} \div \textbf{CancelCar})) \rhd \textbf{ReserveTrain});$$
$$(\textbf{ReserveHotel1} \rhd \textbf{ReserveHotel2})$$

In this transaction, a travel agent wishes to book transportation and a hotel room for a customer. The customer prefers to travel by plane and rental car. These reservations can be canceled. If a flight or rental car is not available, then the agent will book a train ticket to the destination. Once made, the train reservation cannot be canceled. There are two hotel choices at the destination. The first hotel may be full and a reservation may fail, but the second hotel reservation is always successful. We can model the failure and compensation properties of these services as follows:

$\Gamma_t = \langle \textbf{BookFlight} \mapsto \{\checkmark, \times\}, \textbf{CancelFlight} \mapsto \{\checkmark\},$
$\quad\quad \textbf{RentCar} \mapsto \{\checkmark, \times\}, \textbf{CancelCar} \mapsto \{\checkmark\},$
$\quad\quad \textbf{ReserveTrain} \mapsto \{\checkmark, \times\},$
$\quad\quad \textbf{ReserveHotel1} \mapsto \{\checkmark, \times\}, \textbf{ReserveHotel2} \mapsto \{\checkmark\}\rangle$
$\mathcal{C} = \{(\textbf{BookFlight}, \textbf{CancelFlight}), (\textbf{RentCar}, \textbf{CancelCar})\}$

The **ReserveTrain** action is a pivot action, as it has no compensation or exception handler. From inspection, we see that the requirements for cancellation semantics are met:

- If **ReserveTrain** is called, then the actions **BookFlight** and **RentCar** have been compensated by **CancelFlight** and **CancelCar**, respectively.
- If **ReserveHotel1** fails, we recover forward by calling the alternate action **ReserveHotel2**, which cannot fail.

We can check this with our verifier using the following specification predicate:

$$\varphi_{t1} = (((\textbf{BookFlight} \wedge \textbf{RentCar}) \vee \textbf{ReserveTrain}) \wedge$$
$$(\textbf{ReserveHotel1} \vee \textbf{ReserveHotel2})) \vee$$
$$\neg(\textbf{BookFlight} \vee \textbf{RentCar} \vee \textbf{ReserveTrain} \vee$$
$$\textbf{ReserveHotel1} \vee \textbf{ReserveHotel2})$$

The process does indeed satisfy this specification. We can use consistency specifications to check stronger properties as well. For example, we can alter the specification predicate to check that the process does not book both the flight/car and the train and that it only books one hotel:

$$\varphi_{t2} = (((\textbf{BookFlight} \wedge \textbf{RentCar}) \oplus \textbf{ReserveTrain}) \wedge$$
$$(\textbf{ReserveHotel1} \oplus \textbf{ReserveHotel2})) \vee$$
$$\neg(\textbf{BookFlight} \vee \textbf{RentCar} \vee \textbf{ReserveTrain} \vee$$
$$\textbf{ReserveHotel1} \vee \textbf{ReserveHotel2})$$

7. RELATED WORK

Flow Composition Languages. Many formalizations of flow composition languages that support composition and compensation have been proposed in the literature [6, 7, 5]. These formalisms such as STaC [6], the saga calculus [5], and compensating CSP [8] formalize process orchestration using atomic actions and composition operations similar to ours. They differ mostly in the features supported (e.g., whether recursion is allowed, whether parallel processes can be synchronized, or whether there are explicit commit mechanisms), in assumptions on atomic actions (whether or not they always succeed), and in the style of the semantics (trace based or operational).

We chose a core language that includes only the features relevant to our exposition. However, we borrowed extensively from the above languages and believe that similar results hold in the other settings. Like the Saga Calculus, we assume that atomic actions can succeed or fail, as this more closely matches the semantics of industrial languages such as BPEL [12]. We support all the composition operators of the Saga Calculus and Compensating CSP, except that we automatically apply compensation in the event of an error, rather than requiring a transaction block. Our sequential and parallel iterations are inspired by STaC's *generalized parallel* operator. However, our core language does not support interruptible recovery of parallel processes or recursion.

Notions of Correctness. The usual notion of correctness is *cancellation semantics* [8]. One can ensure that a process is self-canceling by restricting processes to have a compensation action for each forward action where compensations cannot fail and are independent of any other compensations running in a parallel branch [8]. Although order-independence between compensations is a realistic restriction, requiring a compensation for each action seems

limiting. Verification becomes more involved when this restriction is relaxed. [15] describes an algorithm which checks that an OPERA workflow, potentially containing non-compensatable actions and exception handling, satisfies cancellation in $O(n^2)$ time. In [19], cancellation is checked on collections of interacting processes by creating an atomicity-equivalent abstraction of each process and checking the product of the abstractions.

Set consistency specifications can capture cancellation semantics. In addition, such specifications can model scenarios where self-cancellation is not desired (e.g., the order case study of section 6.1) and can capture stronger requirements than cancellation (e.g., mutually exclusive actions in the travel agency case study of section 6.2).

Other specification approaches have been suggested for composing web services, independent of compensation and transactional issues. For example, [3] proposes *consistency interfaces*, which define, for a method m of the service and result o of calling that method, the methods called by m and their associated results which lead the result o. The specification language for method calls includes union and intersection, thus providing similar capabilities as a set consistency specification. Consistency interfaces do not treat non-execution of an action the same as atomic failure, and there is no compensation. This precludes the use of negation in specifications and the interpretation of satisfying assignments as executions of the process. Our algorithm can be applied to check processes against consistency interfaces.

Finally, temporal logic specifications, frequently used by model checking tools, can also be used for compensating processes. The problem of checking recursive processes against regular sets of traces is undecidable [11]. Thus, model checkers (e.g., XTL for STaC [16]) usually bound the recursion depth. While temporal logic is a more powerful specification language, set consistency can already model many properties of interest and provides a useful, tractable, and expressive formalism between cancellation and full temporal logic.

8. CONCLUSION

Set consistency is a powerful, yet intuitive, approach to model the correctness requirements of long-running transactions. It permits the modeling of strong consistency requirements such as cancellation, mutual exclusion, and implication. For atomic actions, the equivalence of non-execution, failed execution, and compensated execution is built into the semantics of consistency predicates, making the specifications compact.

The set consistency checking problem is co-NP complete, and can be reduced to checking propositional logic validity. Our verifier generates a predicate from a process definition and specification, which can then be passed to an off-the-shelf SAT solver. We demonstrated the usefulness and tractability of set consistency by running our verifier on three case studies: in each case, the consistency checks could be discharged within a second.

Future work. We plan to extend our core language to include boolean variables, whose values can be tracked by the execution predicate. This permits the modeling and verification of correlated paths in our processes. With this, our core language will be expressive enough to represent BPEL processes, using standard boolean abstraction techniques for data. We are also investigating the extension of set consis-

tency to collections of interacting processes interacting via asynchronous messages.

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