Abstract

In this work we formally verify the soundness and precision of a static program analysis that detects conflicts (e.g. data races) in programs with procedures, thread creation and monitors with the Isabelle theorem prover. As common in static program analysis, our program model abstracts guarded branching by nondeterministic branching, but completely interprets the call-/return behavior of procedures, synchronization by monitors, and thread creation. The analysis is based on the observation that all conflicts already occur in a class of particularly restricted schedules. These restricted schedules are suited to constraint-system-based program analysis.

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1 Introduction

Conflicts are a common programming error in parallel programs. A conflict occurs if the same resource is accessed simultaneously by more than one process. Given a program $\pi$ and two sets of control points $U$ and $V$, the analysis problem is to decide whether there is an execution of $\pi$ that simultaneously reaches one control point from $U$ and one from $V$.

In this work, we use a flowgraph-based program model that extends a previously studied model [6] by reentrant monitors. In our model, programs can call recursive procedures, dynamically create new threads and synchronize via reentrant monitors. As usual in static program analysis, our program model abstracts away guarded branching by nondeterministic choice. We use an operational semantics as reference point for the correctness proofs. It models parallel execution by interleaving, i.e. just one thread is executed at any time and context switches may occur after every step. The next step is nondeterministically selected from all threads ready for execution. The analysis is based on a constraint system generated from the flowgraph. From its least solution, one can decide whether control points from $U$ and $V$ are simultaneously reachable or not.

It is notoriously hard to analyze concurrent programs with constraint systems because of the arbitrary fine-grained interleaving. The key idea behind our analysis is to use a restricted scheduling: While the interleaving semantics can switch the context after each step, the restricted scheduling just allows context switches at certain points of a thread’s execution. We can show that each conflict is also reachable under this restricted scheduling. The restricted schedules can be easily analyzed with constraint systems as most of the complexity generated by arbitrary interleaving does no longer occur due to the restrictions. The remaining concurrency effects can be smoothly handled by using the concept of acquisition histories [5].

Related Work In [6] we present a constraint-system-based analysis for programs with thread creation and procedures but without monitors. The abstraction from synchronization is common in this line of research: There are automata-based techniques [1, 2, 3] as well as constraint-system-based techniques [7, 6] to analyze programs with procedures and either parallel calls or thread creation, but without any synchronization. In [5, 4] analysis techniques for interprocedural parallel programs with a fixed number of initial threads and nested locks are presented. These nested locks are not syntactically bound to the program structure, but assumed to be well-nested, that is any unlock statement is required to release the lock that was acquired last by the thread. Moreover, there is no support for reentrant
locks\textsuperscript{1}. We use monitors instead of locks. Monitors are syntactically bound to the program structure and thus well-nestedness is guaranteed statically. Additionally we directly support reentrant monitors. Our model cannot simulate well-nested locks where a lock statement and its corresponding unlock statement may be in different procedures (as in [5, 4]). As common programming languages like Java also use reentrant monitors rather than locks, we believe our model to be useful as well.

**Document structure** This document contains a commented formalization of these ideas as a collection of Isabelle/HOL theories. A more abstract description is in preparation. This document starts with formalization monitor consistent interleaving (Section 2) and acquisition histories (Section 3). Labeled transition systems are formalized in Section 4, and Section 5 defines the notion of interleaving semantics. Flowgraphs are defined in Section 6, and Section 7 describes their operational semantics. Section 8 contains the formalization of the restricted interleaving and Section 9 contains the constraint systems. Finally, the main result of this development – the correctness of the constraint systems w.r.t. the operational semantics – is briefly stated in Section 10.

## 2 Monitor Consistent Interleaving

**theory** ConsInterleave

**imports** Interleave Misc

**begin**

The monitor consistent interleaving operator is defined on two lists of arbitrary elements, provided an abstraction function \( \alpha \) that maps list elements to pairs of sets of monitors is available. \( \alpha e = (M, M') \) intuitively means that step \( e \) enters the monitors in \( M \) and passes (enters and leaves) the monitors in \( M' \). The consistent interleaving describes all interleavings of the two lists that are consistent w.r.t. the monitor usage.

### 2.1 Monitors of lists of monitor pairs

The following defines the set of all monitors that occur in a list of pairs of monitors. This definition is used in the following context: \( \text{mon-pl} \ (\text{map} \ \alpha \ w) \) is the set of monitors used by a word \( w \) w.r.t. the abstraction \( \alpha \)

**definition**

\[
\text{mon-pl} \ w \equiv \text{foldl} \ (\text{op} \ \cup) \ \{\} \ (\text{map} \ (\lambda e. \ \text{fst} \ e \ \cup \ \text{snd} \ e) \ \ w)
\]

**lemma** \( \text{mon-pl-empty[simp]}: \text{mon-pl} \ [] = \{\} \)

\textsuperscript{1}Reentrant locks can always be simulated by non-reentrant ones, at the cost of a worst-case exponential blowup of the program size
proof
lemma mon-pl-cons[simp]: mon-pl (e#w) = fst e ∪ snd e ∪ mon-pl w
proof
lemma mon-pl-unconc: !!b. mon-pl (a@b) = mon-pl a ∪ mon-pl b
proof
lemma mon-pl-ileq: w ≲ w' ⇒ mon-pl w ⊆ mon-pl w'
proof
lemma mon-pl-set: mon-pl w = ∪ { fst e ∪ snd e | e ∈ set w }
proof
fun cil :: 'a list ⇒ ('a ⇒ ('m set × 'm set)) ⇒ 'a list set
where
— Interleaving with the empty word results in the empty word
[] ⊗α w = {w}
| w ⊗α [] = {w}
— If both words are not empty, we can take the first step of one word, interleave
the rest with the other word and then append the first step to all result set elements,
provided it does not allocate a monitor that is used by the other word
| e1#w1 ⊗α e2#w2 = (if fst (α e1) ∩ mon-pl (map α (e2#w2)) = {} then
e1·(w1 ⊗α e2#w2)
else {} ) ∪ ( if fst (α e2) ∩ mon-pl (map α (e1#w1)) = {} then
e2·(e1#w1 ⊗α w2)
else {} )
| !!w'. | w = e1#w' ; w' ∈ (w1 ⊗α e2#w2);
|        | fst (α e1) ∩ mon-pl (map α (e2#w2)) = {} ] ⇒ P;
| !!w'. | w = e2#w' ; w' ∈ (e1#w1 ⊗α w2);
|        | fst (α e2) ∩ mon-pl (map α (e1#w1)) = {} ] ⇒ P
] ⇒ P
(proof)

lemma cil-last-case-split[cases set, case-names left right]:
| w ∈ e1#w1 ⊗α e2#w2;
| !!w'. | w = e1#w' ; w' ∈ (w1 ⊗α e2#w2);
|        | fst (α e1) ∩ mon-pl (map α (e2#w2)) = {} ] ⇒ P;
| !!w'. | w = e2#w' ; w' ∈ (e1#w1 ⊗α w2);
|        | fst (α e2) ∩ mon-pl (map α (e1#w1)) = {} ] ⇒ P
| ] ⇒ P
(proof)

lemma cil-cases[cases set, case-names both-empty left-empty right-empty app-left
app-right]:


\[ \begin{align*}
&\mbox{cil-contains-empty} \quad | \quad \text{rule-format, simp}: \quad (\epsilon \in wa \otimes_\alpha wb) = (wa=[\ ] \land wb=[\ ]) \\
&\mbox{cil-contains-empty} \quad | \quad \text{rule-format, simp}: \quad (\epsilon \in wa \otimes_\alpha wb) = (wa=[\ ] \land wb=[\ ]) \\
&\mbox{cil-induct-fix} \quad | \quad \text{case-names both-empty left-empty right-empty append}: \quad |
\end{align*} \]
lemma cil-cons-cases[cases set, case-names left right]; [ e#w ∈ w1\otimes_\alpha w2; !!w1'. [w1=e#w1'; w∈w1\otimes_\alpha w2; fst (\alpha e) \cap mon-pl (map \alpha w2) = \{\} ] \implies P; !!w2'. [w2=e#w2'; w∈w1\otimes_\alpha w2'; fst (\alpha e) \cap mon-pl (map \alpha w1) = \{\} ] \implies P ] \implies P
(proof)

lemma cil-set-induct[induct set, case-names empty left right]; !!\alpha w1 w2. [ w∈w1\otimes_\alpha w2; !!\alpha. P [] \alpha [] ]; !!\alpha e w' w1' w2. [ w'∈w1'\otimes_\alpha w2; fst (\alpha e) \cap mon-pl (map \alpha w2) = \{\}; P w' \alpha w1' w2 ] \implies P (e#w') \alpha (e#w1') w2; !!\alpha e w' w2' w1. [ w'∈w1\otimes_\alpha w2'; fst (\alpha e) \cap mon-pl (map \alpha w1) = \{]; P w' \alpha w1 w2' ] \implies P (e#w') \alpha w1 (e#w2')
(proof)

lemma cil-set-induct-fixa[induct set, case-names empty left right]; !!w1 w2. [ w∈w1\otimes_\alpha w2; P [] \alpha [] ]; !!e w' w1' w2. [ w'∈w1'\otimes_\alpha w2; fst (\alpha e) \cap mon-pl (map \alpha w2) = \{\}; P w' \alpha w1' w2 ] \implies P (e#w') \alpha (e#w1') w2; !!e w' w2' w1. [ w'∈w1\otimes_\alpha w2'; fst (\alpha e) \cap mon-pl (map \alpha w1) = \{]; P w' \alpha w1 w2' ] \implies P (e#w') \alpha w1 (e#w2')
(proof)

lemma cil-cons1; [w∈wa\otimes_\alpha wb; fst (\alpha e) \cap mon-pl (map \alpha wb) = \{\}] \implies e#w ∈ e#wa \otimes_\alpha wb
(proof)

lemma cil-cons2; [w∈wa\otimes_\alpha wb; fst (\alpha e) \cap mon-pl (map \alpha wa) = \{\}] \implies e#w ∈ wa \otimes_\alpha e#wb
(proof)

2.2 Properties of consistent interleaving

  — Consistent interleaving is a restriction of interleaving

lemma cil-subset-il; w\otimes_\alpha w' ⊆ w\otimes w'
(proof)

lemma cil-subset-il'; w∈w1\otimes_\alpha w2 \implies w∈w1\otimes w2
(proof)

lemma cil-set; w∈w1\otimes_\alpha w2 \implies set w = set w1 \cup set w2
(proof)

corollary cil-mon-pl; w∈w1\otimes_\alpha w2
\implies mon-pl (map \alpha w) = mon-pl (map \alpha w1) \cup mon-pl (map \alpha w2)
(proof)

lemma cil-length[rule-format]; \forall w∈wa\otimes_\alpha wb. length w = length wa + length wb
(proof)
lemma cil-ileq: \( w \in w_1 \otimes_\alpha w_2 \implies w_1 \preceq w \land w_2 \preceq w \)  
\( \langle \text{proof} \rangle \)

lemma cil-commute: \( w \otimes_\alpha w' = w' \otimes_\alpha w \)  
\( \langle \text{proof} \rangle \)

lemma cil-assoc1: \( !wl w_1 w_2 w_3. \left[ w \in w_1 \otimes_\alpha w_3; w l \in w_1 \otimes_\alpha w_2 \right] \implies \exists wr. w \in w_1 \otimes_\alpha wr \land wr \in w_2 \otimes_\alpha w_3 \)  
\( \langle \text{proof} \rangle \)

lemma cil-assoc2: 
\text{assumes} A: \( w \in w_1 \otimes_\alpha wr \) \text{ and } B: \( wr \in w_2 \otimes_\alpha w_3 \)  
\text{shows} \exists wl. w \in wl \otimes_\alpha w_3 \land wl \in w_1 \otimes_\alpha w_2  
\langle \text{proof} \rangle 

lemma cil-map: \( w \in w_1 \otimes (\alpha \circ f) w_2 \implies \text{map} f w \in \text{map} f w_1 \otimes_\alpha \text{map} f w_2 \)  
\( \langle \text{proof} \rangle \)

end

3 Acquisition Histories

theory AcquisitionHistory  
imports ConsInterleave  
begin

The concept of acquisition histories was introduced by Kahlon, Ivancic, and Gupta [5] as a bounded size abstraction of executions that acquire and release locks that contains enough information to decide consistent interleavability. In this work, we use this concept for reentrant monitors. As in Section 2, we encode monitor usage information in pairs of sets of monitors, and regard lists of such pairs as (abstract) executions. An item \( (E, U) \) of such a list describes a sequence of steps of the concrete execution that first enters the monitors in \( E \) and then passes through the monitors in \( U \). The monitors in \( E \) are never left by the execution. Note that due to the syntactic binding of monitors to the program structure, any execution of a single thread can be abstracted to a sequence of \( (E, U) \)-pairs. Restricting the possible schedules (see Section 8) will allow us to also abstract executions reaching a single program point to a sequence of such pairs.

We want to decide whether two executions are interleavable. The key observation of [5] is, that two executions \( e \) and \( e' \) are not interleavable if and only if there is a conflicting pair \( (m, m') \) of monitors, such that \( e \) enters (and never leaves) \( m \) and then uses \( m' \) and \( e' \) enters (and never leaves) \( m' \) and then uses \( m \).

An acquisition history is a map from monitors to set of monitors. The
acquisition history of an execution maps a monitor $m$ that is allocated at the end of the execution to all monitors that are used after or in the same step that finally enters $m$. Monitors that are not allocated at the end of an execution are mapped to the empty set. Though originally used for a setting without reentrant monitors, acquisition histories also work for our setting with reentrant monitors.

This theory contains the definition of acquisition histories and acquisition history interleavability, an ordering on acquisition histories that reflects the blocking potential of acquisition histories, and a mapping function from paths to acquisition histories that is shown to be compatible with monitor consistent interleaving.

3.1 Definitions

Acquisition histories are modeled as functions from monitors to sets of monitors. Intuitively $m' \in h m$ models that an execution finally is in $m$, and monitor $m'$ has been used (i.e. passed or entered) after or at the same time $m$ has been finally entered. By convention, we have $m \in h m$ or $h m = \{\}$. 

**definition** $ah = \{ (h::'m \Rightarrow 'm set) . \forall m. h m = \{\} \vee m \in h m \}$

**lemma** $ah$-cases[cases set]: $[h\in ah; h m = \{\} \Rightarrow P ; m \in h m \Rightarrow P] \Rightarrow P$

3.2 Interleavability

Two acquisition histories $h1$ and $h2$ are considered interleavable, iff there is no conflicting pair of monitors $m1$ and $m2$, where a pair of monitors $m1$ and $m2$ is called conflicting iff $m1$ is used in $h2$ after entering $m2$ and, vice versa, $m2$ is used in $h1$ after entering $m1$.

**definition** $ah-il ::= ('m \Rightarrow 'm set) \Rightarrow ('m \Rightarrow 'm set) \Rightarrow bool$ (infix $[*] 65$)

**where**

$h1 [*] h2 == \neg(\exists m1 m2. m1 \in h2 m2 \wedge m2 \in h1 m1)$

From our convention, it follows (as expected) that the sets of entered monitors (lock-sets) of two interleavable acquisition histories are disjoint

**lemma** $ah-il$-lockset-disjoint:

$[h1 \in ah; h2 \in ah; h1 [*] h2 ] \Rightarrow h1 m = \{\} \vee h2 m = \{\}$

(proof)

Of course, acquisition history interleavability is commutative

**lemma** $ah-il$-commute: $h1 [*] h2 \Rightarrow h2 [*] h1$

(proof)
3.3 Used monitors

Let’s define the monitors of an acquisition history, as all monitors that occur in the acquisition history

**definition**

\[ \text{mon-ah} :: (\backslash m \Rightarrow \backslash m \text{ set}) \Rightarrow \backslash m \text{ set} \]

\[ \text{mon-ah} \ h \ = \ \bigcup \{ h(m) \mid m. \text{True} \} \]

3.4 Ordering

The element-wise subset-ordering on acquisition histories intuitively reflects the blocking potential: The bigger the acquisition history, the fewer acquisition histories are interleavable with it.

Note that the Isabelle standard library automatically lifts the subset ordering to functions, so we need no explicit definition here.

— The ordering is compatible with interleavability, i.e. smaller acquisition histories are more likely to be interleavable.

**lemma** ah-leq-il:

\[ [ h_1 \ast h_2; h_1' \leq h_1; h_2' \leq h_2 ] \implies h_1' [\ast] h_2' \]

(\text{proof})

**lemma** ah-leq-il-left:

\[ [ h_1 [\ast] h_2; h_1' \leq h_1 ] \implies h_1' [\ast] h_2 \text{ and} \]

ah-leq-il-right:

\[ [ h_1 [\ast] h_2; h_2' \leq h_2 ] \implies h_1 [\ast] h_2' \]

(\text{proof})

3.5 Acquisition histories of executions

Next we define a function that abstracts from executions (lists of enter/use pairs) to acquisition histories

**primrec** ah-ah :: (\backslash m \text{ set} \times \backslash m \text{ set}) list \Rightarrow \backslash m \Rightarrow \backslash m \text{ set} where

\[ \text{ah-ah} \ [] \ m = \{} \]

\[ | \text{ah-ah} \ (e \# w) \ m = (i f m \in \text{fst} \ e \text{ then} \text{fst} \ e \cup \text{snd} \ e \cup \text{mon-pl} \ w \text{ else} \text{ah-ah} \ w \ m) \]

— ah-ah generates valid acquisition histories

**lemma** ah-ah: ah-ah \ w \in ah

(\text{proof})

**lemma** ah-ah-hd:

\[ [ m \in \text{fst} \ e; x \in \text{fst} \ e \cup \text{snd} \ e \cup \text{mon-pl} \ w ] \implies x \in \text{ah-ah} \ (e \# w) \ m \]

(\text{proof})

**lemma** ah-ah-tl:

\[ [ m \notin \text{fst} \ e; x \in \text{ah-ah} \ w \ m ] \implies x \in \text{ah-ah} \ (e \# w) \ m \]

(\text{proof})

**lemma** ah-ah-cases[cases set, case-names hd tl]:

\[ x \in \text{ah-ah} \ w \ m; \]

\[ \text{!!} \ e \ w . [ w = e \# w'; m \in \text{fst} \ e; x \in \text{fst} \ e \cup \text{snd} \ e \cup \text{mon-pl} \ w' ] \implies P; \]

\[ \text{!!} \ e \ w . [ w = e \# w'; m \notin \text{fst} \ e; x \in \text{ah-ah} \ w' \ m ] \implies P \]

\[ \] \implies P
lemma αah-cons-cases [cases set, case-names hd tl]: [x∈α (e#w') m; [m∈fst e; x∈fst e ∪ snd e ∪ mon-pl w'] ⊳ P; [m∉fst e; x∈α w' m] ⊳ P] ⊳ P

lemma mon-ah-subset: mon-ah (αah w) ⊆ mon-pl w

lemma αah-ileq: w1≤w2 ⇒ αah w1 ≤ αah w2

We can now prove the relation of monitor consistent interleavability and interleavability of the acquisition histories.

lemma ah-interleavable1:
  w ∈ w1 ⊗ α w2 ⇒ α (map α w1) [*] α (map α w2)
  — The lemma is shown by induction on the structure of the monitor consistent interleaving operator

lemma ah-interleavable2:
  assumes A: α (map α w1) [*] α (map α w2)
  shows w1 ⊗ α w2 ≠ { }
  — This lemma is shown by induction on the sum of the word lengths

Finally, we can state the relationship between monitor consistent interleaving and interleaving of acquisition histories

theorem ah-interleavable:
  (α (map α w1) [*] α (map α w2)) ←→ (w1⊗αw2≠{ })

3.6 Acquisition history backward update

We define a function to update an acquisition history backwards. This function is useful for constructing acquisition histories in backward constraint systems.

definition ah-update :: (′m ⇒ ′m set) ⇒ (′m set * ′m set) ⇒ ′m set ⇒ (′m ⇒ ′m set)
  where
  ah-update h F M m == if m∈fst F then fst F ∪ snd F ∪ M else h m

Intuitively, ah-update h (E, U) M m means to prepend a step (E, U) to the acquisition history h of a path that uses monitors M. Note that we need the extra parameter M, since an acquisition history does not contain information
about the monitors that are used on a path before the first monitor that
will not be left has been entered.

lemma ah-update-cons: \( \alpha ah (e \# w) = ah-update (\alpha ah w) e \) (mon-pl w)
(proof)

The backward-update function is monotonic in the first and third argument
as well as in the used monitors of the second argument. Note that it is, in
general, not monotonic in the entered monitors of the second argument.

lemma ah-update-mono: \[ \[ h \leq h'; F = F'; M \subseteq M' \] \] \Rightarrow \ ah-update h F M \leq ah-update h' F' M'
(proof)

lemma ah-update-mono2: \[ \[ h \leq h'; U \subseteq U'; M \subseteq M' \] \] \Rightarrow \ ah-update h (E, U) M \leq ah-update h' (E, U') M'
(proof)

end

4 Labeled transition systems

theory LTS
imports Main
begin

Labeled transition systems (LTS) provide a model of a state transition sys-
tem with named transitions.

4.1 Definitions

An LTS is modeled as a ternary relation between start configuration, trans-
ition label and end configuration

type-synonym \((c',a)\) LTS = \((c \times a \times c)\) set

Transitive reflexive closure

inductive-set
\( trcl :: \((c',a)\) LTS \Rightarrow \((c',a\) list)\) LTS
for t
where
empty[simp]: \((c,[]), c\) \in trcl t
| cons[simp]: \[(c,a,c') \in t; (c',w,c'') \in trcl t \] \Rightarrow \( (c,a\#w,c'') \in trcl t \)

4.2 Basic properties of transitive reflexive closure

lemma trcl-empty-cons: \((c,[]), c\) \in trcl t \Rightarrow \(c=c'\)
(proof)

lemma trcl-empty-simp[simp]: \((c,[]), c\) \in trcl t = \(c=c'\)
(proof)
\textbf{lemma} \texttt{trcl-single[simp]}: \((c,[a],c') \in \text{trcl } t\) = \((c,a,c') \in t\)
\textit{(proof)}

\textbf{lemma} \texttt{trcl-uncons}: \((c,a\#w,c')\in\text{trcl } t\) \implies \exists \, ch. \ (c,a,ch)\in t \land (ch,w,c') \in \text{trcl } t
\textit{(proof)}

\textbf{lemma} \texttt{trcl-uncons-cases}:
\[
(c,a\#w,c')\in\text{trcl } S;
\]
\[
!!ch. \ [(c,e,ch)\in S; \ (ch,w,c')\in\text{trcl } S] \implies P
\]
\textit{(proof)}

\textbf{lemma} \texttt{trcl-one-elem}:
\[
(c,e,c')\in t \implies (c,[e],c')\in\text{trcl } t
\]
\textit{(proof)}

\textbf{lemma} \texttt{trcl-unconsE[cases set, case-names split]}:
\[
(c,e\#w,c')\in\text{trcl } S;
\]
\[
!!ch. \ [(c,e,ch)\in S; \ (ch,w,c')\in\text{trcl } S] \implies P
\]
\textit{(proof)}

\textbf{lemma} \texttt{trcl-pair-unconsE[cases set, case-names split]}:
\[
((s,c),e\#w,(s',c'))\in\text{trcl } S;
\]
\[
!!sh \ ch. \ [((s,c),e,(sh,\text{ch}))\in S; \ ((\text{sh},ch),w,(s',c'))\in\text{trcl } S] \implies P
\]
\textit{(proof)}

\textbf{lemma} \texttt{trcl-concat}:
\[
!!c. \ [(c,w1,c')\in\text{trcl } t; \ (c',w2,c'')\in\text{trcl } t] \implies (c,w1@w2,c'')\in\text{trcl } t
\]
\textit{(proof)}

\textbf{lemma} \texttt{trcl-unconcat}:
\[
!!c. \ (c,w1@w2,c')\in\text{trcl } t \implies \exists \, ch. \ (c,w1,\text{ch})\in\text{trcl } t \land (\text{ch},w2,c')\in\text{trcl } t
\]
\textit{(proof)}

\section{4.2.1 Appending of elements to paths}

\textbf{lemma} \texttt{trcl-rev-cons}:
\[
(c,w,\text{ch})\in\text{trcl } T; \ (ch,e,c')\in T \implies (c,w@e,c')\in\text{trcl } T
\]
\textit{(proof)}

\textbf{lemma} \texttt{trcl-rev-uncons}:
\[
(c,w@e,c')\in\text{trcl } T \implies \exists \, ch. \ (c,w,\text{ch})\in\text{trcl } T \land (\text{ch},e,c')\in T
\]
\textit{(proof)}

\textbf{lemma} \texttt{trcl-rev-induct[induct set, consumes 1, case-names empty snoc]}:
\[
!!c'. \ [ (c,w,c')\in\text{trcl } S; \ !!c. \ P \ c \ [] \ c; \ !!c w c' e c''. \ [(c,w,c')\in\text{trcl } S; \ (c',e,c'')\in S; \ P \ c \ w \ c'] \implies P \ c \ (w@e) \ c''
\]
\textit{(proof)}

\textbf{lemma} \texttt{trcl-rev-cases}:
\[
!!c \ c'. \ [
\]
\[
(c,w,c')\in\text{trcl } S; \ [w=\mathbf{[]}; \ c=c']\implies P;
\]
\[
!!ch \ e \ wh. \ [(w=wh@e] ; \ (c,wh,\text{ch})\in\text{trcl } S; \ (ch,e,c')\in S \implies P
\]
\textit{(proof)}
4.2.2 Transitivity reasoning setup

\textbf{declare} \textit{trcl-cons2}\{\textit{trans}\} — It’s important that this is declared before \textit{trcl-concat}, because we want \textit{trcl-concat} to be tried first by the transitivity reasoner

\textbf{declare} \textit{cons}\{\textit{trans}\}
\textbf{declare} \textit{trcl-concat}\{\textit{trans}\}
\textbf{declare} \textit{trcl-rev-cons}\{\textit{trans}\}

4.2.3 Monotonicity

\textbf{lemma} \textit{trcl-mono}: \ (!A B. A \subseteq B \implies trcl A \subseteq trcl B)

\textbf{lemma} \textit{trcl-inter-mono}: \[x \in trcl (S \cap R) \implies x \in trcl S \land x \in trcl (S \cap R) \implies x \in trcl R\]

4.2.4 Special lemmas for reasoning about states that are pairs

\textbf{lemmas} \textit{trcl-pair-induct} = \textit{trcl-induct}[of \ (xc1,xc2) \ xb \ (xa1,xa2), split-format (complete), consumes 1, case-names empty cons]

\textbf{lemmas} \textit{trcl-rev-pair-induct} = \textit{trcl-rev-induct}[of \ (xc1,xc2) \ xb \ (xa1,xa2), split-format (complete), consumes 1, case-names empty snoc]

4.2.5 Invariants

\textbf{lemma} \textit{trcl-prop-trans}\{cases set, consumes 1, case-names empty steps\}: 
\[\[\[\] \implies P; \[\] \implies P; \[\] \implies P\]\]

end

5 Thread Tracking

\textbf{theory} \textit{ThreadTracking}
\textbf{imports} \textit{Main} \sim\sim/\textit{src/HOL/Library/Multiset LTS Misc}
\textbf{begin}

This theory defines some general notion of an interleaving semantics. It defines how to extend a semantics specified on a single thread and a context
to a semantic on multisets of threads. The context is needed in order to keep track of synchronization.

5.1 Semantic on multiset configuration

The interleaving semantics is defined on a multiset of stacks. The thread to make the next step is nondeterministically chosen from all threads ready to make steps.

**definition**

\[
\text{gtr} \text{ gtrs} = \{ \{ \#s\# \} + e, \{ \#s'\# \} + c' \mid s \in \text{cases set, case-names loc other}; \ (s,e, (s',ce') \in \text{gtrs} \}
\]

**lemma** gtr1-s: \((s,c), (s',c') \in \text{gtrs} \implies \{ \#s\# \} + c, \{ \#s'\# \} + c' \in \text{gtr gtrs}

**proof**

**lemma** gtrI: \((s,c), (s',c') \in \text{trcl gtrs} \implies \{ \#s\# \} + c, \{ \#s'\# \} + c' \in \text{trcl (gtr gtrs)}

**proof**

**lemma** gtrE: \[
\langle c, e, c' \rangle \in \text{gtr T};
\]

\[
\text{!!s ce s' ce'.} \ [ c = \{ \#s\# \} + ce; c' = \{ \#s'\# \} + ce' \} ((s,ce), e, (s',ce') \in T \] \implies P
\]

**proof**

**lemma** gtr-empty-conf-s[simp]: \[
\{ \# \} , w, c' \in \text{gtr S}
\]

**proof**

**lemma** gtr-empty-conf1[simp]: \[
\langle \{ \# \}, w, c' \rangle \in \text{trcl (gtr S)} \leftrightarrow (w = \emptyset \land c' = \{ \# \})
\]

**proof**

**lemma** gtr-empty-conf2[simp]: \[
\langle c, w, \{ \# \} \rangle \in \text{trcl (gtr S)} \leftrightarrow (w = \emptyset \land c = \{ \# \})
\]

**proof**

**lemma** gtr-find-thread: \[
\langle c, e, c' \rangle \in \text{gtr gtrs};
\]

\[
\text{!!s ce s' ce'.} \ [ c = \{ \#s\# \} + ce; c' = \{ \#s'\# \} + ce' \} ((s,ce), e, (s',ce') \in \text{gtrs} \] \implies P
\]

**proof**

**lemma** gtr-step-cases[cases set, case-names loc other]: \[
\{ \#s\# \} + c, e, c' \in \text{gtr gtrs};
\]

\[
\text{!!s' ce'.} \ [ c = \{ \#s'\# \} + ce'; (s,ce), e, (s',ce') \in \text{gtrs} \] \implies P;
\]

\[
\text{!!cc ss ss' ce'.} \ [ ce = \{ \#ss\# \} + cc; c' = \{ \#ss'\# \} + ce';
\]

\[
((s, \{ \#s\# \} + cc), e, (s', ce') \in \text{gtrs} \] \implies P
\]

**proof**

**lemma** gtr-rev-cases[cases set, case-names loc other]: \[
\{ c, e, \{ \#s'\# \} + ce' \} \in \text{gtr gtrs};
\]

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5.2 Invariants

**Lemma gtr-preserve-s:**

\[
(c,e,c') \in \text{gtr } T; \quad P c; \quad !s \ c \ s' \ c' \ e. \ [P (\{\#s\#\} + c); ((s,c),e,(s',c')) \in T] \implies P (\{\#s'\#\} + c')
\]

\[\implies P c'\]

\[(proof)\]

**Lemma gtr-preserve:**

\[
(c,w,c') \in \text{trcl } (\text{gtr } T); \quad P c; \quad !s \ c \ s' \ c' \ e. \ [P (\{\#s\#\} + c); ((s,c),e,(s',c')) \in T] \implies P (\{\#s'\#\} + c')
\]

\[\implies P c'\]

\[(proof)\]

5.3 Context preservation assumption

We now assume that the original semantics does not modify threads in the context, i.e. it may only add new threads to the context and use the context to obtain monitor information, but not change any existing thread in the context. This assumption is valid for our semantics, where the context is just needed to determine the set of allocated monitors. It allows us to generally derive some further properties of such semantics.

**Locle env-no-step =**

**Fixes gtrs : \((s \times s \ \text{multiset}) \times \text{LTS}\)**

**Assumes env-no-step-s[cases set, case-names csp]:**

\[
[(s,c),e,(s',c')] \in \text{gtrs}; \quad !\text{csp. } c' \leftarrow \text{csp} + c \implies P \] \implies P
\]

— The property of not changing existing threads in the context transfers to paths

**Lemma (in env-no-step) env-no-step[cases set, case-names csp]:**

\[
((s,c),w,(s',c')) \in \text{trcl gtrs}; \quad !\text{csp. } c' \leftarrow \text{csp} + c \implies P
\]

\[\implies P\]

\[(proof)\]

The following lemma can be used to make a case distinction how a step operated on a given thread in the end configuration:

**Loc** The thread made the step

**Spawn** The thread was spawned by the step
env The thread was not involved in the step

**lemma (in env-no-step) rev-cases-p[cases set, case-names loc spawn env]:**

**assumes STEP: (c,e,\{#s’\#\}+ce)∈gtr gtrs and**

**LOC: \!\! s ce. [ c={#s\#}+ce; (s,ce,e,(s’,ce’))∈gtrs ] ⇒ P and**

**SPAWN: \!\! ss ss’ ce csp.**

\[
\begin{align*}
\text{loc } \Rightarrow P \text{ and}
\end{align*}
\]

**ENV: \!\! ss ss’ ce csp.**

\[
\begin{align*}
\text{env } \Rightarrow P
\end{align*}
\]

shows P

**⟨proof⟩**

### 5.4 Explicit local context

In the multiset semantics, a single thread has no identity. This may become a problem when reasoning about a fixed thread during an execution. For example, in our constraint-system-based approach the operational characterization of the least solution of the constraint system requires to state properties of the steps of the initial thread in some execution. With the multiset semantics, we are unable to identify those steps among all steps.

There are many solutions to this problem, for example, using thread ids either as part of the thread’s configuration or as part of the whole configuration by using lists of stacks or maps from ids to stacks as configuration datatype.

In the following we present a special solution that is strong enough to suit our purposes but not meant as a general solution.

Instead of identifying every single thread uniquely, we only distinguish one thread as the local thread. The other threads are environment threads. We then attach to every step the information whether it was on the local or on some environment thread.

We call this semantics loc/env-semantics in contrast to the multiset-semantics of the last section.

#### 5.4.1 Lifted step datatype

**datatype 'a el-step = LOC 'a | ENV 'a**

**definition**

\[
\begin{align*}
\text{loc } w \text{ == } \text{filter} \ (λ e. \text{case } e \text{ of } \text{LOC } a \Rightarrow \text{True} \ | \ \text{ENV } a \Rightarrow \text{False}) \ w
\end{align*}
\]

**definition**

\[
\begin{align*}
\text{env } w \text{ == } \text{filter} \ (λ e. \text{case } e \text{ of } \text{LOC } a \Rightarrow \text{False} \ | \ \text{ENV } a \Rightarrow \text{True}) \ w
\end{align*}
\]
definition
le-rem-s e == case e of LOC a ⇒ a | ENV a ⇒ a

Standard simplification lemmas
lemma loc-env-simps[simp]:
  loc [] = []
  env [] = []
  ⟨proof⟩

lemma loc-single[simp]: loc [a] = (case a of LOC e ⇒ [a] | ENV e ⇒ [])
  ⟨proof⟩
lemma loc-uncons[simp]:
  loc (a#b) = (case a of LOC e ⇒ [a] | ENV e ⇒ []) @ loc b
  ⟨proof⟩
lemma loc-unconc[simp]: loc (a@b) = loc a @ loc b
  ⟨proof⟩
lemma env-single[simp]: env [a] = (case a of LOC e ⇒ [] | ENV e ⇒ [a])
  ⟨proof⟩
lemma env-uncons[simp]:
  env (a#b) = (case a of LOC e ⇒ [] | ENV e ⇒ [a]) @ env b
  ⟨proof⟩
lemma env-unconc[simp]: env (a@b) = env a @ env b
  ⟨proof⟩

The following simplification lemmas are for converting between paths of the
multiset- and loc/env-semantics
lemma le-rem-simps [simp]:
  le-rem-s (LOC a) = a
  le-rem-s (ENV a) = a
  ⟨proof⟩
lemma le-rem-id-simps[simp]:
  le-rem-loc-LOC = id
  le-rem-loc-ENV = id
  ⟨proof⟩
lemma le-rem-id-map[simp]:
  map le-rem-s (map LOC w) = w
  map le-rem-s (map ENV w) = w
  ⟨proof⟩
lemma env-map-env [simp]: env (map ENV w) = map ENV w
  ⟨proof⟩
lemma env-map-loc [simp]: env (map LOC w) = []
  ⟨proof⟩
lemma loc-map-env [simp]: loc (map ENV w) = []
  ⟨proof⟩
lemma loc-map-loc [simp]: loc (map LOC w) = map LOC w
  ⟨proof⟩
5.4.2  Definition of the loc/env-semantics

type-synonym \( s \text{ el-conf} = (s \times s \text{ multiset}) \)

inductive-set

\[
gtrp :: (s \text{ el-conf}, l) \text{ LTS} \Rightarrow (s \text{ el-conf}, l \text{ el-step}) \text{ LTS}
\]

for \( S \)

where

\[
gtrp\text{-loc}: ((s,c),e,(s',c'))\in S \implies ((s,c),\text{LOC } e,(s',c'))\in gtrp S
\]

\[
gtrp\text{-env}: ((s,\{s\#\}+c),e,(s',\{s\#\}+c'))\in S
\]

\[
\implies ((sl,\{s\#\}+c),\text{ENV } e,(sl,\{s\#\}+c'))\in gtrp S
\]

5.4.3  Relation between multiset- and loc/env-semantics

lemma \( gtrp2gtr\)-s:

\[
(s,c),e,(s',c'))\in gtrp T \implies ((\{s\#\}+c),\text{le-rem-s } e, (\{s\#\}+c')\in gtrp T
\]

\( \langle \text{proof} \rangle \)

lemma \( gtrp2gtr\):

\[
((s,c),w,(s',c'))\in\text{trcl } (gtrp T)
\]

\[
\implies ((\{s\#\}+c),\text{map } \text{le-rem-s } w, (\{s\#\}+c')\in\text{trcl } (gtrp T)
\]

\( \langle \text{proof} \rangle \)

lemma (in env-no-step) \( gtr2gtrp\)-s[cases set, case-names gtrp]:

assumes \( A: ((\{s\#\}+c),e,(s',c'))\in gtr gtrs \)

and CASE: \(!s' ce' ee. [c'=\{s\#\}+ce'; e=\text{le-rem-s } ee; ((s,c),ee,(s',ce'))\in gtr gtrs] \)

shows \( P \)

(\( \langle \text{proof} \rangle \))

lemma (in env-no-step) \( gtr2gtr\)[cases set, case-names gtrp]:

assumes \( A: ((\{s\#\}+c),w,(s',c'))\in\text{trcl } (gtr gtrs) \)

and CASE: \(!s' ce' ee. [c'=\{s\#\}+ce'; w=\text{map } \text{le-rem-s } w w; ((s,c),ww,(s',ce'))\in\text{trcl } (gtr gtrs)] \)

shows \( P \)

(\( \langle \text{proof} \rangle \))

5.4.4  Invariants

lemma \( gtrp\text{-preserve-s} \):

assumes \( A: ((s,c),e,(s',c'))\in gtrp T \)

and INIT: \( P ((\{s\#\}+c) \}

and PRES: \(!s c s' c' e. [P ((\{s\#\}+c): ((s,c),e,(s',c'))\in T] \)

shows \( P ((\{s\#\}+c') \)

(\( \langle \text{proof} \rangle \))
lemma gtrp-preserve:
assumes A: \((s,c),w,(s',c')\)∈trcl (gtrp T)
and INIT: \(P (\{#s\#\}+c)\)
and PRES: \(!!s c s' c' e. [P (\{#s\#\}+c); ((s,c),e,(s',c'))\in T]\)
shows \(P (\{#s'\#\}+c')\)
⟨proof⟩
end

6 Flowgraphs

theory Flowgraph
imports Main Misc
begin

We use a flowgraph-based program model that extends the one we used previously [6]. A program is represented as an edge annotated graph and a set of procedures. The nodes of the graph are partitioned by the procedures, i.e. every node belongs to exactly one procedure. There are no edges between nodes of different procedures. Every procedure has a distinguished entry and return node and a set of monitors it synchronizes on. Additionally, the program has a distinguished main procedure. The edges are annotated with statements. A statement is either a base statement, a procedure call or a thread creation (spawn). Procedure calls and thread creations refer to the called procedure or to the initial procedure of the spawned thread, respectively.

We require that the main procedure and any initial procedure of a spawned thread does not to synchronize on any monitors. This avoids that spawning of a procedure together with entering a monitor is available in our model as an atomic step, which would be an unrealistic assumption for practical problems. Technically, our model would become strictly more powerful without this assumption.

If we allowed this, our model would become strictly more powerful,

6.1 Definitions

datatype ('p,'ba) edgeAnnot = Base 'ba | Call 'p | Spawn 'p
type-synonym ('n,'p,'ba) edge = ('n × ('p,'ba) edgeAnnot × 'n)

record ('n,'p,'ba,'m) flowgraph-rec =
edges :: ('n,'p,'ba) edge set — Set of annotated edges
main :: 'p — Main procedure
entry :: 'p ⇒ 'n — Maps a procedure to its entry point

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\[ \text{return} :: 'p \Rightarrow 'n — \text{Maps a procedure to its return point} \]
\[ \text{mon} :: 'p \Rightarrow 'm \text{ set} — \text{Maps procedures to the set of monitors they allocate} \]
\[ \text{proc-of} :: 'n \Rightarrow 'p — \text{Maps a node to the procedure it is contained in} \]

**definition**
\[
\text{initialproc } \text{fg} \ p = \text{p} = \text{main} \text{ fg} \lor (\exists \ u \ v. (\ u, \text{spawn} \ p, v) \in \text{edges} \ \text{fg})
\]

**lemma** \text{main-is-initial}[\text{simp}]: \text{initialproc } \text{fg} \ (\text{main} \ \text{fg})

\[\langle \text{proof} \rangle\]

**locale** \text{flowgraph} =
\[\text{fixes } \text{fg} :: ('n,'p,'ba,'m,'more) \text{ flowgraph-rec-scheme (structure)}\]

— Edges are inside procedures only
\[\text{assumes edges-part}: (u,a,v) \in \text{edges} \ \text{fg} \implies \text{proc-of} \ \text{fg} \ u = \text{proc-of} \ \text{fg} \ v\]

— The entry point of a procedure must be in that procedure
\[\text{assumes entry-valid}[\text{simp}]: \text{proc-of} \ \text{fg} \ (\text{entry} \ \text{fg} \ p) = p\]

— The return point of a procedure must be in that procedure
\[\text{assumes return-valid}[\text{simp}]: \text{proc-of} \ \text{fg} \ (\text{return} \ \text{fg} \ p) = p\]

— Initial procedures do not synchronize on any monitors
\[\text{assumes initial-no-mon}[\text{simp}]: \text{initialproc} \ \text{fg} \ p \implies \text{mon} \ \text{fg} \ p = \{\}\]

### 6.2 Basic properties

**lemma** (in \text{flowgraph}) \text{spawn-no-mon}[\text{simp}]:
\[
(u, \text{spawn} \ p, v) \in \text{edges} \ \text{fg} \implies \text{mon} \ \text{fg} \ p = \{\}
\]

\[\langle \text{proof} \rangle\]

**lemma** (in \text{flowgraph}) \text{main-no-mon}[\text{simp}]:
\[
\text{mon} \ \text{fg} \ (\text{main} \ \text{fg}) = \{\}
\]

\[\langle \text{proof} \rangle\]

**lemma** (in \text{flowgraph}) \text{entry-return-same-proc}[\text{simp}]:
\[
\text{entry} \ \text{fg} \ p = \text{return} \ \text{fg} \ p' \implies p=p'
\]

\[\langle \text{proof} \rangle\]

**lemma** (in \text{flowgraph}) \text{entry-entry-same-proc}[\text{simp}]:
\[
\text{entry} \ \text{fg} \ p = \text{entry} \ \text{fg} \ p' \implies p=p'
\]

\[\langle \text{proof} \rangle\]

**lemma** (in \text{flowgraph}) \text{return-return-same-proc}[\text{simp}]:
\[
\text{return} \ \text{fg} \ p = \text{return} \ \text{fg} \ p' \implies p=p'
\]

\[\langle \text{proof} \rangle\]

### 6.3 Extra assumptions for flowgraphs

In order to simplify the definition of our restricted schedules (cf. Section 8), we make some extra constraints on flowgraphs. Note that these are no real restrictions, as we can always rewrite flowgraphs to match these constraints, preserving the set of conflicts. We leave it to future work to consider such a rewriting formally.
The background of this restrictions is that we want to start an execution of a thread with a procedure call that never returns. This will allow easier technical treatment in Section 8. Here we enforce this semantic restrictions by syntactic properties of the flowgraph.

The return node of a procedure is called *isolated*, if it has no incoming edges and is different from the entry node. A procedure with an isolated return node will never return. See Section 8.1 for a proof of this.

**definition**
\[
\text{isolated-ret \ fg \ p} \equiv \\
(\forall u \ l. \neg(u,l,\text{return \ fg \ p}) \in \text{edges \ fg}) \land \text{entry \ fg \ p} \neq \text{return \ fg \ p}
\]

The following syntactic restrictions guarantee that each thread’s execution starts with a non-returning call. See Section 8.1 for a proof of this.

**locale** eflowgraph = flowgraph +
— Initial procedure’s entry node isn’t equal to its return node
**assumes** initial-no-ret: initialproc \ fg \ p \implies \text{entry \ fg \ p} \neq \text{return \ fg \ p}
— The only outgoing edges of initial procedures’ entry nodes are call edges to procedures with isolated return node
**assumes** initial-call-no-ret: \ [\ [\ \text{initialproc} \ \fg \ \p; \ (\text{entry} \ \fg \ \p, \ l, \ v) \in \text{edges} \ \fg] \] \implies \exists \p'. \ l=\text{Call} \ \p' \land \text{isolated-ret} \ \fg \ \p'

### 6.4 Example Flowgraph

This section contains a check that there exists a (non-trivial) flowgraph, i.e. that the assumptions made in the flowgraph and eflowgraph locales are consistent and have at least one non-trivial model.

**definition**
\[
\text{example-fg} \equiv \{\\n\text{edges} = \{(0::nat,0::nat),\text{Call} \ 1,(0,1)), ((1,0),\text{Spawn} \ 0,(1,0)), ((1,0),\text{Call} \ 0, (1,0))\},\\n\text{main} = 0,\\n\text{entry} = \lambda p. \ (p,0),\\n\text{return} = \lambda p. \ (p,1),\\n\text{mon} = \lambda p. \text{if} \ p=1 \text{ then} \{0\} \text{ else} \{\},\\n\text{proc-of} = \lambda (p,x). \ p \}\\n\]

**lemma** exists-eflowgraph: eflowgraph example-fg
\{\text{proof}\}

### 7 Operational Semantics

**theory** Semantics
**imports** Main Flowgraph ~/src/HOL/Library/Multiset LTS Interleave ThreadTracking

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begin

7.1 Configurations and labels

The state of a single thread is described by a stack of control nodes. The top node is the current control node and the nodes deeper in the stack are stored return addresses. The configuration of a whole program is described by a multiset of stacks.

Note that we model stacks as lists here, the first element being the top element.

type-synonym 'n conf = ('n list) multiset

A step is labeled according to the executed edge. Additionally, we introduce a label for a procedure return step, that has no corresponding edge.

datatype (′p,′ba) label = LBase 'ba | LCall 'p | LRet | LSpawn 'p

7.2 Monitors

The following defines the monitors of nodes, stacks, configurations, step labels and paths (sequences of step labels)

definition — The monitors of a node are the monitors the procedure of the node synchronizes on
mon-n fg n == mon fg (proc-of fg n)

definition — The monitors of a stack are the monitors of all its nodes
mon-s fg s == \{ mon-n fg n | n . n \in set s \}

definition — The monitors of a configuration are the monitors of all its stacks
mon-c fg c == \{ mon-s fg s | s . s :# c \}

definition — The monitors of a step label are the monitors of procedures that are called by this step
mon-e :: (′b,′c,′d,′a,′e) flowgraph-rec-scheme ⇒ (′c,′f) label ⇒ ′a

lemma mon-e-simps [simp]:
mon-e fg (LBase a) = {}
mon-e fg (LCall p) = mon fg p
mon-e fg (LRet) = {}
mon-e fg (LSpawn p) = {}
(proof)
definition
mon-w fg w == \{ mon-e fg e | e . e \in set w \}
lemma mon-s-alt: mon-s fg s == \( \bigcup (\text{mon fg \ ' proc-of fg \ ' set s}) \)
⟨proof⟩

lemma mon-c-alt: mon-c fg c == \( \bigcup (\text{mon fg \ ' set-of c}) \)
⟨proof⟩

lemma mon-w-alt: mon-w fg w == \( \bigcup (\text{mon-c fg \ ' set w}) \)
⟨proof⟩

lemma mon-sI: \[ n \in \text{set s}; m \in \text{mon-n fg n} \] \( \Rightarrow \) m \( \in \) mon-s fg s
⟨proof⟩

lemma mon-sD: m \( \in \) mon-s fg s \( \Rightarrow \) \exists n \in \text{set s}. m \( \in \) mon-n fg n
⟨proof⟩

lemma mon-n-same-proc: proc-of fg n = proc-of fg n’ \( \Rightarrow \) mon-n fg n = mon-n fg n’
⟨proof⟩

lemma mon-s-same-proc: proc-of fg \ ' set s = proc-of fg \ ' set s’ \( \Rightarrow \) mon-s fg s = mon-s fg s’
⟨proof⟩

lemma (in flowgraph) mon-of-entry[simp]: mon-n fg (entry fg p) = mon fg p
⟨proof⟩

lemma (in flowgraph) mon-of-ret[simp]: mon-n fg (return fg p) = mon fg p
⟨proof⟩

lemma mon-c-single[simp]: mon-c fg \{\#\} = mon-s fg s
⟨proof⟩

lemma mon-s-single[simp]: mon-s fg \[n\] = mon-n fg n
⟨proof⟩

lemma mon-s-empty[simp]: mon-s fg [] = {}
⟨proof⟩

lemma mon-c-empty[simp]: mon-c fg \# = {}
⟨proof⟩

lemma mon-s-unconc: mon-s fg (a@b) = mon-s fg a \cup mon-s fg b
⟨proof⟩

lemma mon-s-uncons[simp]: mon-s fg (a#as) = mon-n fg a \cup mon-s fg as
⟨proof⟩

lemma mon-c-unconc: mon-c fg (a+b) = mon-c fg a \cup mon-c fg b
⟨proof⟩

lemma mon-cl: \[ s:#c; m \in mon-s fg s \] \( \Rightarrow \) m \( \in \) mon-c fg c
⟨proof⟩

lemma mon-cD: \[ m \in mon-c fg c \] \( \Rightarrow \) \exists s. s:#c \wedge m \in mon-s fg s
⟨proof⟩

lemma mon-s-mono: set s \subseteq set s’ \( \Rightarrow \) mon-s fg s \subseteq mon-s fg s’
⟨proof⟩
lemma mon-c-mono: \( c \leq c' \implies \text{mon-c \ fg \ c} \subseteq \text{mon-c \ fg \ c}' \)

\begin{proof}
\end{proof}

lemma mon-w-empty[simp]: \( \text{mon-w \ fg \ []} = \{\} \)

\begin{proof}
\end{proof}

lemma mon-w-single[simp]: \( \text{mon-w \ fg \ [e]} = \text{mon-e \ fg \ e} \)

\begin{proof}
\end{proof}

lemma mon-w-unconc: \( \text{mon-w \ fg \ (wa@wb)} = \text{mon-w \ fg \ wa} \cup \text{mon-w \ fg \ wb} \)

\begin{proof}
\end{proof}

lemma mon-w-uncons[simp]: \( \text{mon-w \ fg \ (e#w)} = \text{mon-e \ fg \ e} \cup \text{mon-w \ fg \ w} \)

\begin{proof}
\end{proof}

lemma mon-w-ileq: \( w \preceq w' \implies \text{mon-w \ fg \ w} \subseteq \text{mon-w \ fg \ w}' \)

\begin{proof}
\end{proof}

7.3 Valid configurations

We call a configuration valid if each monitor is owned by at most one thread.

definition valid fg c == \( \forall s s'. \{\#s\} + \{\#s'\} \leq c \implies \text{mon-s \ fg \ s} \cap \text{mon-s \ fg \ s}' = \{} \)

lemma valid-empty[simp, intro!]: \( \text{valid \ fg \ \{\}\} \)

\begin{proof}
\end{proof}

lemma valid-single[simp, intro!]: \( \text{valid \ fg \ \{s\}\} \)

\begin{proof}
\end{proof}

lemma valid-split1: \( \text{valid \ fg \ (c+c')} \implies \text{valid \ fg \ c} \land \text{valid \ fg \ c'} \land \text{mon-c \ fg \ c} \cap \text{mon-c \ fg \ c}' = \{} \)

\begin{proof}
\end{proof}

lemma valid-split2: \( [\text{valid \ fg \ c}; \text{valid \ fg \ c'}; \text{mon-c \ fg \ c} \cap \text{mon-c \ fg \ c}' = \{\}] \implies \text{valid \ fg \ (c+c')} \)

\begin{proof}
\end{proof}

lemma valid-unconc: \( \text{valid \ fg \ (c+c')} \longleftrightarrow (\text{valid \ fg \ c} \land \text{valid \ fg \ c'} \land \text{mon-c \ fg \ c} \cap \text{mon-c \ fg \ c}' = \{\}) \)

\begin{proof}
\end{proof}

lemma valid-no-mon: \( \text{mon-c \ fg \ c} = \{} \implies \text{valid \ fg \ c} \)

\begin{proof}
\end{proof}

7.4 Configurations at control points

— A stack is at U if its top node is from the set U

primrec atU-s :: 'n set => 'n list => bool where

\( \text{atU-s \ U \ []} = \text{False} \)

\( | \text{atU-s \ U \ (u#r)} = (u \in U) \)

lemma atU-s-decomp[simp]: \( \text{atU-s \ U \ (s@ss')} = (\text{atU-s \ U \ s} \lor (s=[] \land \text{atU-s \ U \ s'})) \)

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\textbf{definition} \\
$atU \ U \ c \ == \ \exists \ s. \ s:\# \ c \ \land \ atU-s \ U \ s$ \\

\textbf{lemma} $atU-fmt$: $[\ atU \ U \ c; \ !u \ r. \ [u\#r \ : \ # \ c; \ ui\in U] \ \Longrightarrow \ P] \ \Longrightarrow \ P$ \\
\textbf{proof} \\

\textbf{lemma} $atU-union-cases[case-names \ left \ right, \ consumes \ 1]$: \\
$\begin{array}{l}
atU \ U \ (c1+c2); \\
\atU \ U \ c1 \ \Longrightarrow \ P; \\
\atU \ U \ c2 \ \Longrightarrow \ P \\
\end{array}$ \\
$\Longrightarrow \ P$ \\
\textbf{proof} \\

\textbf{lemma} $atU-add$: $atU \ U \ c \ \Longrightarrow \ atU \ U \ (c+ce) \ \land \ atU \ U \ (ce+c)$ \\
\textbf{proof} \\

\textbf{lemma} $atU-empty[simp]$: $\neg atU \ U \ {\#}$ \\
\textbf{proof} \\

\textbf{lemma} $atU-single[simp]$: $atU \ U \ {\#s\#} = atU-s \ U \ s$ \\
\textbf{proof} \\

\textbf{lemma} $atU-single-top[simp]$: $atU \ U \ {\#u\#r\#} = (u\in U)$ \\
\textbf{proof} \\

\textbf{lemma} $atU-xchange-stack$: $atU \ U \ ({\#u\#r\#}+c) \ \Longrightarrow \ atU \ U \ ({\#u\#r'\#}+c)$ \\
\textbf{proof} \\

\textbf{definition} \\
$atUV \ U \ V \ c \ == \ \exists \ su \ sv. \ \{\#\} + \{\#sv\} \ \leq \ c \ \land \ atU-s \ U \ su \ \land \ atU-s \ V \ sv$ \\

\textbf{lemma} $atUV-empty[simp]$: $\neg atUV \ U \ V \ {\#}$ \\
\textbf{proof} \\

\textbf{lemma} $atUV-single[simp]$: $\neg atUV \ U \ V \ {\#s\#}$ \\
\textbf{proof} \\

\textbf{lemma} $atUV-add$: $atUV \ U \ V \ (c1+c2) \leftrightarrow \ (atUV \ U \ V \ c1) \ \lor \ (atUV \ U \ V \ c2) \ \lor \ (atU \ U \ c1 \ \land \ atU \ V \ c2) \ \lor \ (atU \ V \ c1 \ \land \ atU \ U \ c2)$ \\
\textbf{proof} \\

\textbf{lemma} $atUV-union-cases[case-names \ left \ right \ lr \ rl, \ consumes \ 1]$: \\
$atUV \ U \ V \ (c1+c2)$; \\

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atUV U V c1 \Rightarrow P;
atUV U V c2 \Rightarrow P;\[ \[ \atUV U c1; \atUV V c2 \] \Rightarrow P;\[ \[ \atUV V c1; \atUV U c2 \] \Rightarrow P \]
\langle proof \rangle

7.5 Operational semantics

7.5.1 Semantic reference point

We now define our semantic reference point. We assess correctness and completeness of analyses relative to this reference point.

\textbf{inductive-set}
\[ \text{refpoint} :: (\text{'}n,\text{'p,}ba,\text{'}m,\text{'more}) \text{ flowgraph-rec-scheme} \Rightarrow \text{'}n \text{ conf} \times (\text{'}p,ba) \text{ label} \times \text{'}n \text{ conf} \text{ set} \]
\textbf{for fg}
\textbf{where}

— A base edge transforms the top node of one stack and leaves the other stacks untouched.
\[ \text{refpoint-base}: \[ (u,\text{Base} a,v) \in \text{edges } fg; \text{ valid } fg (\{\#u\#r\#\}+c) \] \Rightarrow (\{\#u\#r\#\}+c,\text{LBase} a,\{\#v\#r\#\}+c) \in \text{refpoint } fg \]

— A call edge transforms the top node of a stack and then pushes the entry node of the called procedure onto that stack. It can only be executed if all monitors the called procedure synchronizes on are available. Reentrant monitors are modeled here by checking availability of monitors just against the other stacks, not against the stack of the thread that executes the call. The other stacks are left untouched.
\[ \text{refpoint-call}: \[ (u,\text{Call} p,v) \in \text{edges } fg; \text{ valid } fg (\{\#u\#r\#\}+c); \text{ mon } fg p \cap \text{ mon-c } fg c = \{} \] \Rightarrow (\{\#u\#r\#\}+c,\text{LCall} p,\{\#\text{entry } fg p,v\#r\#\}+c) \in \text{refpoint } fg \]

— A return step pops a return node from a stack. There is no corresponding flowgraph edge for a return step. The other stacks are left untouched.
\[ \text{refpoint-ret}: \[ \text{ valid } fg (\{\#\text{return } fg p\#r\#\}+c) \] \Rightarrow (\{\#\text{return } fg p\#r\#\}+c,\text{LRet}(\{\#r\#\}+c) \in \text{refpoint } fg \]

— A spawn edge transforms the top node of a stack and adds a new stack to the environment, with the entry node of the spawned procedure at the top and no stored return addresses. The other stacks are also left untouched.
\[ \text{refpoint-spawn}: \[ (u,\text{Spawn} p,v) \in \text{edges } fg; \text{ valid } fg (\{\#u\#r\#\}+c) \] \Rightarrow (\{\#u\#r\#\}+c,\text{LSpawn} p,\{\#v\#r\#\}+\{\#\text{entry } fg p\#\}+c) \in \text{refpoint } fg \]

Instead of working directly with the reference point semantics, we define the operational semantics of flowgraphs by describing how a single stack is transformed in a context of environment threads, and then use the theory developed in Section 5 to derive an interleaving semantics. Note that this semantics is also defined for invalid configurations (cf. Section 7.3). In Section 7.6.1 we will show that it preserves validity of a configuration, and in Section 7.6.2 we show that it is equivalent to the reference point semantics
on valid configurations.

**inductive-set**

\[
trss :: (\text{\textquoteleft}n,\text{\textquoteleft}p,\text{\textquoteleft}ba,\text{\textquoteleft}m,\text{\textquoteleft}more\text{\textquoteleft} flowgraph-rec-scheme \Rightarrow
\]

\[\text{\textquoteleft}(n \text{ list} \ast \text{\textquoteleft}n \text{ conf} \ast \text{\textquoteleft}p,\text{\textquoteleft}ba \ast \text{\textquoteleft}n \text{ list} \ast \text{\textquoteleft}n \text{ conf}) \text{ set}
\]

**for fg**

**where**

\[
trss-base: \begin{array}{l}
\left[ (u,\text{Base} \ a, v) \in \text{edges} \ fg \right] \Rightarrow \\
\ \ \ \ ((u \# r, c), \text{LBase} \ a, (v \# r, c)) \in trss \ fg
\end{array}
\]

\[
\begin{array}{l}
trss-call: \begin{array}{l}
\left[ (u,\text{Call} \ p, v) \in \text{edges} \ fg; \text{mon} \ fg \ p \cap \text{mon-c} \ fg \ c = \{\}\right] \Rightarrow \\
\ \ \ \ ((u \# r, c), \text{LCall} \ p, ((\text{entry} \ fg \ p) \# v \# r, c)) \in trss \ fg
\end{array}
\end{array}
\]

\[
\begin{array}{l}
trss-ret: \begin{array}{l}
((\text{return} \ fg \ p) \# r, c), \text{LRet} \ (r, c) \in trss \ fg
\end{array}
\end{array}
\]

\[
\begin{array}{l}
trss-spawn: \begin{array}{l}
\left[ (u, \text{Spawn} \ p, v) \in \text{edges} \ fg \right] \Rightarrow \\
\ \ \ \ ((u \# r, c), \text{LSpawn} \ p, (v \# r, ([\text{entry} \ fg \ p] \# r) + c)) \in trss \ fg
\end{array}
\end{array}
\]

— The interleaving semantics is generated using the general techniques from Section 5

**abbreviation** tr where tr fg == gtr (trss fg)

— We also generate the loc/env-semantics

**abbreviation** trp where trp fg == gtrp (trss fg)

### 7.6 Basic properties

#### 7.6.1 Validity

**lemma** (in flowgraph) trss-valid-preserve-s:

\[
\begin{array}{l}
\text{valid fg} ((\# s \#) + c); ((s, c), e, (s, c')) \in trss \ fg \Rightarrow \text{valid fg} ((\# s \#) + c')
\end{array}
\]

(proof)

**lemma** (in flowgraph) trss-valid-preserve:

\[
\begin{array}{l}
\left[ ((s, c), w, (s, c')) \in \text{trcl} \ (trss \ fg); \text{valid fg} ((\# s \#) + c) \right] \Rightarrow \text{valid fg} ((\# s \#) + c')
\end{array}
\]

(proof)

**lemma** (in flowgraph) tr-valid-preserve-s:

\[
\begin{array}{l}
\left[ (c, e, c') \in \text{tr} \ fg; \text{valid fg} \ c \right] \Rightarrow \text{valid fg} \ c'
\end{array}
\]

(proof)

**lemma** (in flowgraph) tr-valid-preserve:

\[
\begin{array}{l}
\left[ (c, w, e, c') \in \text{trcl} \ (tr \ fg); \text{valid fg} \ c \right] \Rightarrow \text{valid fg} \ c'
\end{array}
\]

(proof)

**lemma** (in flowgraph) trp-valid-preserve-s:

\[
\begin{array}{l}
\left[ ((s, c), e, (s, c')) \in \text{trp} \ fg; \text{valid fg} ((\# s \#) + c) \right] \Rightarrow \text{valid fg} ((\# s \#) + c')
\end{array}
\]

(proof)

**lemma** (in flowgraph) trp-valid-preserve:

\[
\begin{array}{l}
\left[ ((s, c), w, (s, c')) \in \text{trcl} \ (trp \ fg); \text{valid fg} ((\# s \#) + c) \right] \Rightarrow \text{valid fg} ((\# s \#) + c')
\end{array}
\]

(proof)
7.6.2 Equivalence to reference point

— The equivalence between the semantics that we derived using the techniques from Section 5 and the semantic reference point is shown nearly automatically.

**Lemma** `refpoint-eq-s`: `valid fg c \implies ((c, e, c') \in refpoint fg) \iff ((c, e, c') \in tr fg)`

**Lemma** *(in flowgraph)\* `refpoint-eq`:

`valid fg c \implies ((c, w, c') \in trcl (refpoint fg)) \iff ((c, w, c') \in trcl (tr fg))`

**(proof)**

7.6.3 Case distinctions

**Lemma** `trss-c-cases-s[cases set, case-names no-spawn spawn]`: \[
((s, c), e, (s', c')) \in trss fg;
\[ c' = c \] \implies P;
\[ !p u v. e = LSpawn p; (u, Spawn p, v) \in edges fg;
\[ \text{hd } s = u; \text{hd } s' = v; c' = \text{[\# entry } fg p \text{ ]} + c \] \implies P
\]

**(proof)**

**Lemma** *(in flowgraph)\* `trss-c-fmt-s`:

\[
((s, c), e, (s', c')) \in trss fg \]\[\implies \exists csp. c' = csp + c \land (csp = \{\#\} \lor (\exists p. e = LSpawn p \land csp = \{\# \text{ entry } fg p \} ))\]

**(proof)**

**Lemma** *(in flowgraph)\* `trss-c-split-s`:

\[
((s, c), e, (s', c')) \in trss fg;
\[ !csp. c' = csp + c \land \text{mon-c fg csp} = \{} \] \implies P
\]

**(proof)**

**Lemma** *(in flowgraph)\* `trss-c-cases[cases set, case-names c-case]`: \[!s c. ((s, c), w, (s', c')) \in trcl (trss fg);
\[ !\text{csp}. c' = csp + c \land \text{\[!s. s:\#csp} \implies \exists p u v. s = \text{[entry } fg p \text{ ]} \land (u, Spawn p, v) \in edges fg \land \text{\text{initialproc } fg p } ]\]

\[\] \implies P

**(proof)**

**Lemma** *(in flowgraph)\* `c-of-initial-no-mon`:

**assumes** `A`: \[!s. s:\#csp \implies \exists p. s = \text{[entry } fg p \text{ ]} \land \text{\text{initialproc } fg p }]

**shows** `\text{\text{mon-c } fg csp} = \{\}`

**(proof)**

**Lemma** *(in flowgraph)\* `trss-c-no-mon-s`:

**assumes** `A`: \[((s, c), e, (s', c')) \in trss fg \]

**shows** `\text{\text{mon-c } fg c'} = \text{\text{mon-c } fg c}`

**(proof)**
corollary \([\text{in flowgraph}]\) \(\text{trss-c-no-mon}\):
\[
\{(s,c),w,(s',c')\} \in \text{trcl} (\text{trss } fg) \implies \text{mon } fg \ c' = \text{mon } fg \ c
\]
⟨proof⟩

lemma \([\text{in flowgraph}]\) \(\text{trss-spawn-no-mon-step}[\text{simp}]\):
\[
\{(s,c),L\text{Spawn } p, (s',c')\} \in \text{trss } fg \implies \text{mon } fg \ p = \{\}
\]
⟨proof⟩

lemma \(\text{trss-no-empty-s}[\text{simp}]\):
\[
\{(\{\},c),(s',c')\} \in \text{trss } fg = \text{False}
\]
⟨proof⟩

lemma \(\text{trss-no-empty}[\text{simp}]\):
\[
\text{assumes } A: \{(\{\},c),(s',c')\} \in \text{trcl} (\text{trss } fg)
\text{ shows } w=\{\} \land s'=\{\} \land c=c'
\]
⟨proof⟩

lemma \(\text{trs-step-cases}[\text{cases set}, \text{case-names NO-SPAWN SPAWN}]\):
\[
\text{assumes } A: (c,e,c') \in \text{tr } fg
\text{assumes } A-\text{NO-SPAWN}: \forall s \in \text{csp. } [\]
\[
(s,e),(s',c')\} \in \text{trss } fg;
\begin{align*}
\text{c}=\{\#s\}+ce; \ c'=\{\#s'\}+ce
\end{align*}
\]
\text{assumes } A-\text{SPAWN}: \forall s \in \text{p. } [\]
\[
(s,e),L\text{Spawn } p,(s',\{\#entry \text{fg } p\}+ce)\} \in \text{trss } fg;
\begin{align*}
\text{c}=\{\#s\}+ce; \\
\text{c'}=\{\#s'\}+\{\#entry \text{fg } p\}+ce; \ e=L\text{Spawn } p
\end{align*}
\text{shows } P
\]
⟨proof⟩

7.7 Advanced properties

7.7.1 Stack composition / decomposition

lemma \(\text{trss-stack-comp-s}\):
\[
\{(s,c),e,(s',c')\} \in \text{trss } fg \implies (s@r,c),e,(s'@r,c')\} \in \text{trss } fg
\]
⟨proof⟩

lemma \(\text{trss-stack-comp}\):
\[
\{(s,c),w,(s',c')\} \in \text{trcl} (\text{trss } fg) \implies (s@r,c),w,(s'@r,c')\} \in \text{trcl} (\text{trss } fg)
\]
⟨proof⟩

lemma \(\text{trss-stack-decomp-s}\):
\[
\exists sp'. \ s'=sp'@r \land (s,c),e,(sp',c')\} \in \text{trss } fg
\]
⟨proof⟩
lemma trss-find-return: 
  \[ ((s@r,c),w,(r',c')) \in \text{trcl} (\text{trss } fg); \]
  \[ \not\exists \! w\! w' \! b\! h \! c\! h. \! \begin{array}{l} \! \![w=wa@wb; \![((s,c),wa,([],ch)) \in \text{trcl} (\text{trss } fg)]; \!
  \end{array} \]
  \((r,ch),wb,(r',c')) \in \text{trcl} (\text{trss } fg) \] \implies P

If \( s = [] \), the proposition follows trivially

⟨proof⟩

lemma trss-return-cases[cases set]: \( !u \! r \! c. \]
  \[ ((u\neq r,c),w,(r',c')) \in \text{trcl} (\text{trss } fg); \]
  \[ !w\! w' \! b\! h \! c\! h. \! \begin{array}{l} \! \![w=wa@wb; \![((u,c),wa,([],ch)) \in \text{trcl} (\text{trss } fg)]; \!
  \end{array} \]
  \((r,ch),wb,(r',c')) \in \text{trcl} (\text{trss } fg) \] \implies P

⟨proof⟩

lemma (in flowgraph) trss-find-call:
  \[ !u \! r'\! c'. \!
  \begin{array}{l} \!
  \end{array} \!
  \Rightarrow \exists \! rh \! c\! h \! p \! w\! a\! w\! b \! . \!
  \begin{array}{l} \!
  \end{array} \!
  \[ w=wa@((LCall \! p)\! \# \! wb \wedge \]
  \[ \text{proc-of } fg \! v \! = \! p \wedge \]
  \[ (([sp],c),wa,(rh,\! ch)) \! \in \! \text{trcl} \! (\! \text{trss } fg) \wedge \]
  \[ ((rh,\! ch),LCall \! p,((\text{entry } fg \! p)\! \# \! r'\! c')) \! \in \! \text{trss } fg \wedge \]
  \[ ((\text{entry } fg \! p),\! ch),wb,([v],c')) \! \in \! \text{trcl} \! (\! \text{trss } fg) \]

⟨proof⟩

lemma (in flowgraph) trss-find-call':
  assumes A: \[ (([sp],c),w,(\text{return } fg \! p)\! \# \! [u],c')) \in \text{trcl} \! (\! \text{trss } fg) \]
  and EX: \[ !uh \! c\! h \! w\! a\! w\! b \! . \!
  \begin{array}{l} \!
  \end{array} \!
  \[ w=wa@((LCall \! p)\! \# \! wb); \]
  \[ (([sp],c),wa,(uh,\! ch)) \! \in \! \text{trcl} \! (\! \text{trss } fg); \]
  \[ ((uh,\! ch),LCall \! p,((\text{entry } fg \! p)\! \# \! [u],ch)) \! \in \! \text{trss } fg; \]
  \[ (uh,\! Call \! p,\! u') \! \in \! \text{edges } fg; \]
  \[ ((\text{entry } fg \! p),\! ch),wb,((\text{return } fg \! p),\! c')) \! \in \! \text{trcl} \! (\! \text{trss } fg) \]

shows P

⟨proof⟩

lemma (in flowgraph) trss-bot-proc-const:
  \[ !s'\! u'\! c'. \!
  \begin{array}{l} \!
  \end{array} \!
  \[ ((s@u],c),w,(s'@[u'],c')) \in \text{trcl} \! (\! \text{trss } fg) \]

⟨proof⟩

lemma (in flowgraph) trss-er-path-proc-const:
  \[ ((\text{entry } fg \! p),\! c),w,((\text{return } fg \! q),\! c') \in \text{trcl} \! (\! \text{trss } fg) \] \Rightarrow p=q

⟨proof⟩

lemma trss-2empty-to-2return: \[ (\!(s,c),w,([,c'])) \in \text{trcl} \! (\! \text{trss } fg); \!
  \[ \not\exists \! w' \! p. \!
  \begin{array}{l} \!
  \end{array} \!
  \[ w=w'^{\!@}(\text{LRet}) \wedge ((s,c),w',((\text{return } fg \! p),c')) \! \in \! \text{trcl} \! (\! \text{trss } fg) \]

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lemma \text{trss-2return-to-2empty}: \[ ((s,c),w,([\text{return}\ fg\ p],c')) \in \text{trcl}\ (\text{trss}\ fg) \] \[ \implies ((s,c),w\#[\text{LRet}],[[]],c')) \in \text{trcl}\ (\text{trss}\ fg) \]

(proof)

7.7.2 Adding threads

lemma \text{trss-env-increasing-s}: ((s,c),e,(s',c')) \in \text{trss}\ fg \implies c \leq c'

(proof)

lemma \text{trss-env-increasing}: ((s,c),w,(s',c')) \in \text{trcl}\ (\text{trss}\ fg) \implies c \leq c'

(proof)

7.7.3 Conversion between environment and monitor restrictions

lemma \text{trss-mon-e-no-ctx}:

\[(s,c),e,(s',c') \in \text{trss}\ fg \implies \text{mon-e}\ fg\ e \cap \text{mon-e}\ fg\ c = \{\}\]

(proof)

lemma (in flowgraph) \text{trss-mon-w-no-ctx}:

\[(s,c),w,(s',c') \in \text{trcl}\ (\text{trss}\ fg) \implies \text{mon-w}\ fg\ w \cap \text{mon-e}\ fg\ c = \{\}\]

(proof)

lemma (in flowgraph) \text{trss-modify-context-s}:

\[!\text{cn}.\ [(s,c),e,(s',c') \in \text{trss}\ fg; \text{mon-e}\ fg\ e \cap \text{mon-e}\ fg\ cn = \{\}]] \]

\[\implies \exists\ \text{csp}.\ c' = \text{csp} + c \land \text{mon-e}\ fg\ \text{csp} = \{\} \land ((s,cn),e,(s',\text{csp}+\text{cn})) \in \text{trss}\ fg\]

(proof)

lemma (in flowgraph) \text{trss-modify-context}[rule-format]:

\[[(s,c),w,(s',c') \in \text{trcl}\ (\text{trss}\ fg)]\]

\[\implies \forall\ \text{cn}.\ \text{mon-w}\ fg\ w \cap \text{mon-e}\ fg\ cn = \{\}\]

\[\implies (\exists\ \text{csp}.\ c' = \text{csp} + c \land \text{mon-e}\ fg\ \text{csp} = \{\} \land ((s,cn),w,(s',\text{csp}+\text{cn})) \in \text{trcl}\ (\text{trss}\ fg))\]

(proof)

lemma \text{trss-add-context-s}:

\[[(s,c),e,(s',c') \in \text{trss}\ fg; \text{mon-e}\ fg\ e \cap \text{mon-e}\ fg\ ce = \{\}]] \]

\[\implies ((s,c+ce),e,(s',c'+ce)) \in \text{trss}\ fg\]

(proof)

lemma \text{trss-add-context}:

\[[(s,c),w,(s',c') \in \text{trcl}\ (\text{trss}\ fg); \text{mon-w}\ fg\ w \cap \text{mon-e}\ fg\ ce = \{\}]] \]

\[\implies ((s,c+ce),w,(s',c'+ce)) \in \text{trcl}\ (\text{trss}\ fg)\]

(proof)

lemma \text{trss-drop-context-s}:

\[\[(s,c+ce),e,(s',c'+ce)) \in \text{trss}\ fg\]\n
\[\implies ((s,c),e,(s',c')) \in \text{trss}\ fg\land \text{mon-e}\ fg\ e \cap \text{mon-e}\ fg\ ce = \{\}]]\]

(proof)

lemma \text{trss-drop-context}:

\[!s\ c.\ [(s,c+ce),w,(s',c'+ce)) \in \text{trcl}\ (\text{trss}\ fg)]\]

\[\implies ((s,c),w,(s',c')) \in \text{trcl}\ (\text{trss}\ fg)\land \text{mon-w}\ fg\ w \cap \text{mon-e}\ fg\ ce = \{\}]

(proof)
theory Normalization
imports Main ThreadTracking Semantics ConsInterleave
begin

The idea of normalized paths is to consider particular schedules only. While
the original semantics allows a context switch to occur after every single
step, we now define a semantics that allows context switches only before
non-returning calls or after a thread has reached its final stack. We then
show that this semantics is able to reach the same set of configurations as

8 Normalized Paths

end
the original semantics.

8.1 Semantic properties of restricted flowgraphs

It makes the formalization smoother, if we assume that every thread’s execution begins with a non-returning call. For this purpose, we defined syntactic restrictions on flowgraphs already (cf. Section 6.3). We now show that these restrictions have the desired semantic effect.

— Procedures with isolated return nodes will never return

\[ \text{lemma (in } \text{eflowgraph)} \ ISO-RET-NO-RET: \forall u \ c. \]
\[ ISO-RET \ f g \ p; \]
\[ \text{proc-of } f g \ u = p; \]
\[ u \neq \text{return } f g \ p; \]
\[ (([u],c),w,([\text{return } f g \ p],c')) \in \text{trcl (trss } f g) \]
\[ \Rightarrow \text{False} \]

\[ \langle \text{proof} \rangle \]

\[ \text{lemma (in } \text{eflowgraph)} \ INITIAL-STARTS-WITH-CALL: \]
\[ \bigl(\bigl([\text{entry } f g \ p],c,e,(s',c')\bigr) \in \text{trss } f g; \text{initialproc } f g \ p \bigr) \]
\[ \Rightarrow \exists p', e=\text{LCall } p' \land \text{isolated-ret } f g \ p' \]

\[ \langle \text{proof} \rangle \]

\[ \text{lemma (in } \text{eflowgraph)} \ NO-SL-FROM-INITIAL: \]
\[ \text{assumes } A: w \neq [] \]
\[ \text{initialproc } f g \ p \]
\[ (([\text{entry } f g \ p],c),w,([v],c')) \in \text{trcl (trss } f g) \]
\[ \text{shows False} \]

\[ \langle \text{proof} \rangle \]

\[ \text{lemma (in } \text{eflowgraph)} \ NO-RETSL-FROM-INITIAL: \]
\[ \text{assumes } A: w \neq [] \]
\[ \text{initialproc } f g \ p \]
\[ (([\text{entry } f g \ p],c),w,(r',c')) \in \text{trcl (trss } f g) \]
\[ \text{length } r' \leq 1 \]
\[ \text{shows False} \]

\[ \langle \text{proof} \rangle \]

8.2 Definition of normalized paths

In order to describe the restricted schedules, we define an operational semantics that performs an atomically scheduled sequence of steps in one step, called a \textit{macrostep}. Context switches may occur after macrosteps only. We call this the \textit{normalized semantics} and a sequence of macrosteps a \textit{normalized path}.

Since we ensured that every path starts with a non-returning call, we can define a macrostep as an initial call followed by a same-level path\textsuperscript{2} of the

\textsuperscript{2}Same-level paths are paths with balanced calls and returns. The stack-level at the beginning of their execution is the same as at the end, and during the execution, the stack never falls below the initial level.
called procedure. This has the effect that context switches are either performed before a non-returning call (if the thread makes a further macrostep in the future) or after the thread has reached its final configuration.

As for the original semantics, we first define the normalized semantics on a single thread with a context and then use the theory developed in Section 5 to derive interleaving semantics on multisets and configurations with an explicit local thread (loc/env-semantics, cf. Section 5.4).

**inductive-set**

\[ ntrs :: (\texttt{\_n}\texttt{\_p}\texttt{\_ba}\texttt{\_m}\texttt{\_more}) \texttt{flowgraph-rec-scheme} \Rightarrow \]

\[ ((\texttt{\_n list} \times \texttt{\_n conf}) \times (\texttt{\_p}\texttt{\_ba}) \texttt{label list} \times (\texttt{\_n list} \times \texttt{\_n conf}) \texttt{set} \]

**for fg**

**where**

— A macrostep transforms one thread by first calling a procedure and then doing a same-level path

\[ ntrs-step: (((\texttt{\_u r ce})\texttt{LCall p} \texttt{entry fg p} \# \texttt{u'} \# \texttt{r ce}) \in \texttt{trss fg}; \]

\[ (((\texttt{\_entry fg p} \texttt{r ce})\texttt{w} \texttt{\_v u' r ce'}) \in \texttt{trcl (trss fg)}]] \Rightarrow \]

\[ (((\texttt{\_u r ce})\texttt{LCall p\#w} \texttt{v u' r ce'}) \in \texttt{ntrs fg} \]

**abbreviation** \[ ntr where ntr fg == \texttt{gt} (ntrs fg) \]

**abbreviation** \[ ntrp where ntrp fg == \texttt{gt}p (ntrs fg) \]

**interpretation** \[ ntrs: env-no-step ntrs fg \]

\langle proof\rangle

### 8.3 Representation property for reachable configurations

In this section, we show that a configuration is reachable if and only if it is reachable via a normalized path.

The first direction is to show that a normalized path is also a path. This follows from the definitions. Note that we first show that a single macrostep corresponds to a path and then generalize the result to sequences of macrosteps

**lemma** \[ ntrs-is-trss-s: ((\texttt{s c},\texttt{w s' c'}) \in \texttt{ntrs fg} \Rightarrow ((\texttt{s c},\texttt{w s' c'}) \in \texttt{trcl (trss fg)}} \]

\langle proof\rangle

**lemma** \[ ntrs-is-trss: ((\texttt{s c},\texttt{w s' c'}) \in \texttt{trcl (ntrs fg)}} \]

\[ \Rightarrow ((\texttt{s c},\texttt{foldl (op @) w s' c'}) \in \texttt{trcl (trss fg)}} \]

\langle proof\rangle

**lemma** \[ ntr-is-tr-s: (\texttt{c w c'}) \in \texttt{ntr fg} \Rightarrow (\texttt{c w c'}) \in \texttt{trcl (tr fg)}} \]

\langle proof\rangle

**lemma** \[ ntr-is-tr: (\texttt{c w w' c'}) \in \texttt{trcl (ntr fg)} \Rightarrow (\texttt{c foldl (op @) w w' c'}) \in \texttt{trcl (tr fg)}} \]

\langle proof\rangle
The other direction requires to prove that for each path reaching a configuration there is also a normalized path reaching the same configuration. We need an auxiliary lemma for this proof, that is a kind of append rule: Given a normalized path reaching some configuration \(c\), and a same level or returning path from some stack in \(c\), we can derive a normalized path to \(c\) modified according to the same-level path. We cannot simply append the same-level or returning path as a macrostep, because it does not start with a non-returning call. Instead, we will have to append it to some macrostep in the normalized path, i.e. move it „left” into the normalized path.

Intuitively, we can describe the concept of the proof as follows: Due to the restrictions we made on flowgraphs, a same-level or returning path cannot be the first steps on a thread. Hence there is a last macrostep that was executed on the thread. When this macrostep was executed, all threads held less monitors then they do at the end of the execution, because the set of monitors held by every single thread is increasing during the execution of a normalized path. Thus we can append the same-level or returning path to the last macrostep on that thread. As a same-level or returning path does not allocate any monitors, the following macrosteps remain executable. If we have a same-level path, appending it to a macrostep yields a valid macrostep again and we are done. Appending a returning path to a macrostep yields a same-level path. In this case we inductively repeat our argument.

The actual proof is strictly inductive; it either appends the same-level path to the last macrostep or inductively repeats the argument.

**lemma (in eflowgraph) ntr-sl-move-left:** !ce u r w r′ ce′.

\[
\begin{align*}
&\exists \{\text{entry}\ fg\ p\}\ w,\\{\# \ u\ \#\ \#\}\ +\ ce \in \text{trcl}\ (ntr\ fg); \\
&\{\{u,ce\},w,(r′,ce′)\}\in \text{trcl}\ (\text{trss}\ fg); \\
&\text{initialproc}\ fg\ p; \\
&\text{length}\ r′ \leq 1;\ w\neq[] \\
&\Rightarrow \exists \text{ww′}. (\{\# \text{entry}\ fg\ p\}\ ,\ \text{ww′}\ ,\{\# r′@r\ \#\} + ce′) \in \text{trcl}\ (ntr\ fg)
\end{align*}
\]

(\text{proof})

Finally we can prove: Any reachable configuration can also be reached by a normalized path. With eflowgraph.ntr-sl-move-left we can easily show this. With eflowgraph.ntr-sl-move-left we can easily show this by induction on the reaching path. For the empty path, the proposition follows trivially. Else we consider the last step. If it is a call, we can execute it as a macrostep and get the proposition. Otherwise the last step is a same-level (Base, Spawn) or returning (Ret) path of length 1, and we can append it to the normalized path using eflowgraph.ntr-sl-move-left.

**lemma (in eflowgraph) normalize:**

\[
\begin{align*}
&\text{normalize}: [ \\
&(cstart,w,c′)\in \text{trcl}\ (tr\ fg); \\
&\text{cstart}=\{\#\ \text{entry}\ fg\ p\ \#\}; \\
&\text{initialproc}\ fg\ p ] \\
&\Rightarrow \exists \text{ww′}. (\{\# \text{entry}\ fg\ p\ \#\},w′,c′)\in \text{trcl}\ (ntr\ fg)
\end{align*}
\]

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— The lemma is shown by induction on the reaching path

\langle \text{proof} \rangle

As the main result of this section we get: A configuration is reachable if and only if it is also reachable via a normalized path:

**Theorem (in flowgraph) ntr-repr:**

\[(\exists w. \{\#\text{entry } fg (\text{main } fg)\#\}, w, c) \in \text{trcl } (\text{tr } fg) \leftrightarrow (\exists w. \{\#\text{entry } fg (\text{main } fg)\#\}, w, c) \in \text{trcl } (\text{ntr } fg)\]

\langle \text{proof} \rangle

### 8.4 Properties of normalized path

Like a usual path, also a macrostep modifies one thread, spawns some threads and preserves the state of all the other threads. The spawned threads do not make any steps, thus they stay in their initial configurations.

**Lemma ntrs-c-cases-s[cases set]:**

\[
\begin{align*}
\left((s, c), w, (s', c')\right) &\in \text{ntr } fg; \\
\not\exists \text{sp. } \left(c' = \text{sp} + c; \not\exists s. s: \#csp \implies \exists p u v. s = [\text{entry } fg p] \land (u, \text{Spawn } p, v) \in \text{edges } fg \land \text{initialproc } fg p\right) \\
\| &\implies P
\end{align*}
\]

\langle \text{proof} \rangle

**Lemma ntrs-c-cases[cases set]:**

\[
\begin{align*}
\left((s, c), w, (s', c')\right) &\in \text{trcl } (\text{ntr } fg); \\
\not\exists \text{sp. } \left(c' = \text{sp} + c; \not\exists s. s: \#csp \implies \exists p u v. s = [\text{entry } fg p] \land (u, \text{Spawn } p, v) \in \text{edges } fg \land \text{initialproc } fg p\right) \\
\| &\implies P
\end{align*}
\]

\langle \text{proof} \rangle

### 8.4.1 Validity

Like usual paths, also normalized paths preserve validity of the configurations.

**Lemmas (in flowgraph) ntrs-valid-preserve-s = trss-valid-preserve[OF ntrs-is-trss-s]**

**Lemmas (in flowgraph) ntr-valid-preserve-s = tr-valid-preserve[OF ntr-is-tr-s]**

**Lemmas (in flowgraph) ntrs-valid-preserve = trss-valid-preserve[OF ntrs-is-trss]**

**Lemmas (in flowgraph) ntr-valid-preserve = tr-valid-preserve[OF ntr-is-tr]**

**Lemma (in flowgraph) ntrp-valid-preserve-s:**

**assumes A:** \(((s, c), e, (s', c'))\in \text{ntrp } fg\)

**and V:** valid \(fg (\{\#s\#\} + c)\)

**shows valid \(fg (\{\#s'\#\} + c')\)**

\langle \text{proof} \rangle

**Lemma (in flowgraph) ntrp-valid-preserve:**
assumes $A: ((s, c), e, (s', c')) \in \text{trcl} \ (\text{ntrp fg})$
and $V: \text{valid fg} \ (\{#s#\} + c)$
shows $\text{valid fg} \ (\{#s'#\} + c')$
(proof)

8.4.2 Monitors

The following defines the set of monitors used by a normalized path and shows its basic properties:

definition mon-ww fg $ww = \text{foldl} \ (\text{op} \cup) \ {} \ (\text{map} \ (\text{mon-w fg}) \ \text{ww})$

definition mon-loc fg $ww = \text{mon-ww fg} \ (\text{map le-rem-s} \ (\text{loc \ ww}))$

definition mon-env fg $ww = \text{mon-ww fg} \ (\text{map le-rem-s} \ (\text{env \ ww}))$

lemma mon-ww-empty [simp]: $\text{mon-ww fg} \ [] = {}$
(proof)
lemma mon-ww-uncons [simp]:
$\text{mon-ww fg} \ (\text{ee} \# \ \text{ww}) = \text{mon-w fg} \ \text{ee} \cup \ \text{mon-ww fg} \ \text{ww}$
(proof)
lemma mon-ww-unconc: $\text{mon-ww fg} \ (\text{ww1} @ \text{ww2}) = \text{mon-ww fg} \ \text{ww1} \cup \text{mon-ww fg} \ \text{ww2}$
(proof)

lemma mon-env-empty [simp]: $\text{mon-env fg} \ [] = {}$
(proof)
lemma mon-env-single [simp]:
$\text{mon-env fg} \ [e] = (\text{case } e \ \text{of} \ \text{LOC} \ a \Rightarrow {} \mid \ \text{ENV} \ a \Rightarrow \text{mon-w fg} \ a)$
(proof)
lemma mon-env-uncons [simp]:
$\text{mon-env fg} \ (e\#w) = (\text{case } e \ \text{of} \ \text{LOC} \ a \Rightarrow {} \mid \ \text{ENV} \ a \Rightarrow \text{mon-w fg} \ a) \cup \text{mon-env fg} \ w$
(proof)
lemma mon-env-unconc: $\text{mon-env fg} \ (w1 @ w2) = \text{mon-env fg} \ w1 \cup \text{mon-env fg} \ w2$
(proof)

lemma mon-loc-empty [simp]: $\text{mon-loc fg} \ [] = {}$
(proof)
lemma mon-loc-single [simp]:
$\text{mon-loc fg} \ [e] = (\text{case } e \ \text{of} \ \text{ENV} \ a \Rightarrow {} \mid \ \text{LOC} \ a \Rightarrow \text{mon-w fg} \ a)$
(proof)
lemma mon-loc-uncons [simp]:
$\text{mon-loc fg} \ (e\#w) = (\text{case } e \ \text{of} \ \text{ENV} \ a \Rightarrow {} \mid \ \text{LOC} \ a \Rightarrow \text{mon-w fg} \ a) \cup \text{mon-loc fg} \ w$
(proof)
**lemma** mon-loc-unconc:
\[
\text{mon-loc fg } (w_1 \otimes w_2) = \text{mon-loc fg } w_1 \cup \text{mon-loc fg } w_2
\]
⟨proof⟩

**lemma** mon-ww-of-foldl[simp]: mon-w fg (foldl (op @) [] ww) = mon-ww fg ww
⟨proof⟩

**lemma** mon-ww-ileq: \( w \preceq w' \implies \text{mon-ww fg } w \subseteq \text{mon-ww fg } w' \)
⟨proof⟩

**lemma** mon-ww-cil: \( w \in w_1 \otimes \alpha \otimes w_2 \implies \text{mon-ww fg } w = \text{mon-ww fg } w_1 \cup \text{mon-ww fg } w_2 \)
⟨proof⟩

**lemma** mon-loc-cil: \( w \in w_1 \otimes \alpha \otimes w_2 \implies \text{mon-loc fg } w = \text{mon-loc fg } w_1 \cup \text{mon-loc fg } w_2 \)
⟨proof⟩

**lemma** mon-env-cil: \( w \in w_1 \otimes \alpha \otimes w_2 \implies \text{mon-env fg } w = \text{mon-env fg } w_1 \cup \text{mon-env fg } w_2 \)
⟨proof⟩

**lemma** mon-ww-of-le-rem: \( \text{mon-ww fg } (\text{map le-rem-s } w) = \text{mon-loc fg } w \cup \text{mon-env fg } w \)
⟨proof⟩

**lemma** mon-env-ileq: \( w \preceq w' \implies \text{mon-env fg } w \subseteq \text{mon-env fg } w' \)
⟨proof⟩

**lemma** mon-loc-ileq: \( w \preceq w' \implies \text{mon-loc fg } w \subseteq \text{mon-loc fg } w' \)
⟨proof⟩

**lemma** mon-loc-map-loc[simp]: \( \text{mon-loc fg } (\text{map LOC } w) = \text{mon-ww fg } w \)
⟨proof⟩

**lemma** mon-env-map-env[simp]: \( \text{mon-env fg } (\text{map ENV } w) = \text{mon-ww fg } w \)
⟨proof⟩

**lemma** mon-loc-map-env[simp]: \( \text{mon-loc fg } (\text{map ENV } w) = \{\} \)
⟨proof⟩

**lemma** mon-env-map-loc[simp]: \( \text{mon-env fg } (\text{map LOC } w) = \{\} \)
⟨proof⟩

**lemma** (in flowgraph) ntrs-mon-increasing-s: \((s,c),e,(s',c')\)\(\in\) ntrs fg
\(\implies\) \(\text{mon-s fg } s \subseteq \text{mon-s fg } s' \land \text{mon-c fg } c = \text{mon-c fg } c' \)
⟨proof⟩

**lemma** (in flowgraph) ntr-mon-increasing-s: \((c,ce,c')\)\(\in\) ntr fg
\(\implies\) \(\text{mon-c fg } c \subseteq \text{mon-c fg } c' \)
⟨proof⟩

**lemma** (in flowgraph) ntrp-mon-increasing-s: \((s,c),e,(s',c')\)\(\in\) ntrp fg
\[ \implies \text{mon-s fg } s \subseteq \text{mon-s fg } s' \land \text{mon-c fg } c \subseteq \text{mon-c fg } c' \]

**Lemma (in flowgraph) ntrp-mon-increasing:** \( ((s,c),e, (s',c')) \in \text{trcl (ntrp fg)} \)
\[ \implies \text{mon-s fg } s \subseteq \text{mon-s fg } s' \land \text{mon-c fg } c \subseteq \text{mon-c fg } c' \]

**8.4.3 Modifying the context**

**Lemmas (in flowgraph) ntrs-c-no-mon-s = trss-c-no-mon[OF ntrs-is-trss-s]**

**Lemmas (in flowgraph) ntrs-c-no-mon = trss-c-no-mon[OF ntrs-is-trss]**

Also like a usual path, a normalized step must not use any monitors that are allocated by other threads

**Lemmas (in flowgraph) ntrs-mon-e-no-ctx = trss-mon-w-no-ctx[OF ntrs-is-trss-s]**

**Lemmas (in flowgraph) ntrs-mon-e-no-ctx:**
\[ \text{assumes } A: ((s,c),w, (s',c')) \in \text{trcl (ntrp fg)} \]
\[ \text{shows } \text{mon-w fg } w \cap \text{mon-c fg } c = \{ \} \]

**Lemma (in flowgraph) ntrp-mon-env-e-no-ctx:**
\[ ((s,c),w, (s',c')) \in \text{trcl (ntrp fg)} \]
\[ \implies \text{mon-env fg } w \cap \text{mon-s fg } s = \{ \} \]

**Lemma (in flowgraph) ntrp-mon-loc-e-no-ctx:**
\[ ((s,c),w, (s',c')) \in \text{trcl (ntrp fg)} \]
\[ \implies \text{mon-loc fg } w \cap \text{mon-c fg } c = \{ \} \]

**The next lemmas are rules how to add or remove threads while preserving the executability of a path**

**Lemma (in flowgraph) ntrs-modify-context-s:**
\[ \text{assumes } A: ((s,c),e, (s',c')) \in \text{ntrp fg} \]
\[ \text{and } B: \text{mon-w fg } ee \cap \text{mon-c fg } cn = \{ \} \]
\[ \text{shows } \exists \text{csp. } c' = \text{csp} + c \land \text{mon-c fg } esp = \{ \} \land ((s,cn),e,(s',csp+cn)) \in \text{ntrp fg} \]

**Lemma (in flowgraph) ntrs-modify-context[rule-format]:**
\[ [((s,c),w, (s',c')) \in \text{trcl (ntrp fg)}] \]
\[ \implies \forall cn. \text{mon-ww fg } w \cap \text{mon-c fg } cn = \{ \} \]
\[ \implies (\exists \text{csp. } c' = \text{csp} + c \land \text{mon-c fg } esp = \{ \} \land ((s,cn),w,(s',csp+cn)) \in \text{trcl (ntrp fg)}) \]
proof

lemma ntrs-xchange-context-s:
assumes A: ((s,c),ee,(s',csp+c))∈ntrs fg
and B: mon-c fg cn ⊆ mon-c fg c
shows ((s,cn),ee,(s',csp+cn))∈ntrs fg

lemma ntrs-replace-context-s:
assumes A: ((s,c+cr),ee,(s',c'+cr))∈ntrs fg
and B: mon-c fg crn ⊆ mon-c fg cr
shows ((s,c+crn),ee,(s',c'+crn))∈ntrs fg

lemma (in flowgraph) ntrs-xchange-context: !!s c c' cn. [ ((s,c),ww,(s',c'))∈trcl (ntrs fg);
   mon-c fg cn ⊆ mon-c fg c ] ⇒ \exists csp.
   c'=csp+c ∧ ((s,cn),ww,(s',csp+cn))∈trcl (ntrs fg)

lemma (in flowgraph) ntrs-replace-context:
assumes A: ((s,c+cr),ww,(s',c'+cr))∈trcl (ntrs fg)
and B: mon-c fg crn ⊆ mon-c fg cr
shows ((s,c+crn),ww,(s',c'+crn))∈trcl (ntrs fg)

lemma (in flowgraph) ntr-add-context-s:
assumes A: (c,e,c')∈ntr fg
and B: mon-w fg e ∩ mon-c fg cn = {}
shows (c+cn,e,c'+cn)∈ntr fg

lemma (in flowgraph) ntrp-add-context-s:
[[ (c,w,c)∈trcl (ntr fg); mon-w fg w ∩ mon-c fg cn = {} ]]
⇒ (c+cn,w,c'+cn)∈trcl (ntr fg)

lemma (in flowgraph) ntrp-add-context:
[[ (c,w,c)∈trcl (ntrp fg); mon-w fg w ∩ mon-c fg cn = {} ]]
⇒ (c+cn,w,c'+cn)∈trcl (ntrp fg)

lemma (in flowgraph) ntrp-add-context-s:
assumes A: ((s,c),e,(s',c'))∈ntrp fg
and B: mon-w fg e ∩ mon-c fg cn = {}
shows ((s,c+cn),e,(s',c'+cn))∈ntrp fg

lemma (in flowgraph) ntrs-add-context-s:
assumes A: ((s,c),e,(s',c'))∈ntrs fg
and B: mon-w fg e ∩ mon-c fg cn = {}
shows ((s,c+cn),e,(s',c'+cn))∈ntrs fg

lemma (in flowgraph) ntrs-add-context:
[[ ((s,c),e,(s',c'))∈trcl (ntrs fg); mon-w fg (le-rem-s e) ∩ mon-c fg cn = {} ]]
⇒ ((s,c+cn),e,(s',c'+cn))∈trcl (ntrs fg)
\begin{proof}

\textbf{lemma (in flowgraph) ntrp-add-context:} \[
((s,c),w,(s',c'))\in trcl (ntrp fg); 
\text{mon-ww fg } (\text{map le-rem-s } w) \cap \text{ mon-c fg } cn = \{}
\implies ((s,c+cn),w,(s',c'+cn))\in trcl (ntrp fg)
\]

\end{proof}

8.4.4 Altering the local stack

\textbf{lemma ntrp-stack-comp-s:}
\begin{itemize}
\item \textbf{assumes} A: \(((s,c),ee,(s',c'))\in ntrp fg
\item \textbf{shows} \(((s@r,c),ee,(s'@r,c'))\in ntrp fg
\end{itemize}
\begin{proof}
\end{proof}

\textbf{lemma ntrp-stack-comp:} \(((s,c),ww,(s',c'))\in trcl (ntrp fg) \implies ((s@r,c),ww,(s'@r,c'))\in trcl (ntrp fg)
\begin{proof}
\end{proof}

\textbf{lemma (in flowgraph) ntrp-stack-comp-s:}
\begin{itemize}
\item \textbf{assumes} A: \(((s,c),ee,(s',c'))\in ntrp fg
\item \textbf{and} B: \text{mon-s fg } r \cap \text{ mon-env fg } [ee] = \{}
\item \textbf{shows} \(((s@r,c),ee,(s'@r,c'))\in ntrp fg
\end{itemize}
\begin{proof}
\end{proof}

\textbf{lemma (in flowgraph) ntrp-stack-comp:}
\begin{itemize}
\item \textbf{assumes} A: \(((s,c),ww,(s',c'))\in trcl (ntrp fg); \text{mon-s fg } r \cap \text{ mon-env fg } ww = \{}
\item \textbf{shows} \(((s@r,c),ww,(s'@r,c'))\in trcl (ntrp fg)
\end{itemize}
\begin{proof}
\end{proof}

\textbf{lemma ntrp-stack-top-decomp-s:}
\begin{itemize}
\item \textbf{assumes} A: \(((u#r,c),ee,(s',c'))\in ntrp fg
\item \textbf{and} EX: !u' \ p.
\item \textbf{shows} P
\end{itemize}
\begin{proof}
\end{proof}

\textbf{lemma ntrp-stack-decomp-s:}
\begin{itemize}
\item \textbf{assumes} A: \(((u#s@r,c),ee,(s',c'))\in ntrp fg
\item \textbf{and} EX: !v \ u' \ p.
\item \textbf{shows} P
\end{itemize}
\begin{proof}
\end{proof}

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\textbf{lemma} \ ntrs-stack-decomp: \ \[ \forall u \ s \ r \ c \ P. \]
\[
((u\#s@r,c),ww,(s',c'))\in trcl (ntrs fg);
\]
\[
\forall v \ rr. \ [s'=v\#rr@r; ((u\#s,c),ww,(v\#rr,c'))\in trcl (ntrs fg)] \implies P
\]
\[\] \[\implies P\]
\[\langle \text{proof}\rangle\]

\textbf{lemma} \ ntrp-stack-decomp-s:
\textbf{assumes} \ A: \ ((u\#s@r,c),ee,(s',c'))\in ntrp fg
\textbf{and} \ EX: \ \[s'=v\#rr@r; ((u\#s,c),ee,(v\#rr,c'))\in ntrp fg \] \implies P
\textbf{shows} \ P
\[\langle \text{proof}\rangle\]

\textbf{lemma} \ ntrp-stack-decomp: \ \[\forall u \ s \ r \ c \ P. \]
\[
((u\#s@r,c),ww,(s',c'))\in trcl (ntrs fg);
\]
\[
\forall v \ rr. \ [s'=v\#rr@r; ((u\#s,c),ww,(v\#rr,c'))\in trcl (ntrs fg)] \implies P
\]
\[\] \[\implies P\]
\[\langle \text{proof}\rangle\]

\section{Relation to monitor consistent interleaving}

In this section, we describe the relation of the consistent interleaving operator (cf. Section 2) and the macrostep-semantics.

\subsection{Abstraction function for normalized paths}

We first need to define an abstraction function that maps a macrostep on a pair of entered and passed monitors, as required by the $\otimes_{\alpha}$-operator:

A step on a normalized paths enters the monitors of the first called procedure and passes the monitors that occur in the following same-level path.

\textbf{definition}
\[\alpha_{n\ fg\ e} \ ::= \ \text{if } e=\[] \ \text{then } ((\{\},\{\}) \ \text{else } (\text{mon}-e\ fg\ (hd\ e), \text{mon}-w\ fg\ (tl\ e))\]

\textbf{lemma} \ \alpha_{n\ simps}[simp]:
\[\alpha_{n\ fg}\ [] = ((\{\},\{\})\]
\[\alpha_{n\ fg}\ (e\#w) = (\text{mon}-e\ fg\ e, \text{mon}-w\ fg\ w)\]
\[\langle \text{proof}\rangle\]

\textbf{definition}
\[\alpha_{nl\ fg\ e} \ ::= \ \alpha_{n\ fg}\ (\text{le}-\text{rem}-s\ e)\]

\textbf{lemma} \ \alpha_{nl\ def}': \ \alpha_{nl\ fg} \ ::= \ \alpha_{n\ fg} \circ \text{le}-\text{rem}-s
\[\langle \text{proof}\rangle\]

\textbf{lemma} \ \alpha_{nl\ simps}[simp]:
\[\alpha_{nl\ fg}\ (\text{ENV}\ x) = \alpha_{n\ fg}\ x\]
\[\alpha_{nl\ fg}\ (\text{LOC}\ x) = \alpha_{n\ fg}\ x\]
\[\langle \text{proof}\rangle\]

\textbf{lemma} \ \alpha_{nl\ simps1}[simp]:
\((\alpha n fg) \circ ENV = \alpha n fg\)
\((\alpha n fg) \circ LOC = \alpha n fg\)
\(\langle \text{proof} \rangle\)

**Lemma** \(\alpha n-\alpha nl\): \((\alpha n fg) \circ \text{le-rem-s} = \alpha n fg\)
\(\langle \text{proof} \rangle\)

**Lemma** \(\alpha n-fst-snd\[\text{simp}]\): \(\text{fst} (\alpha n fg w) \cup \text{snd} (\alpha n fg w) = \text{mon-w} fg w\)
\(\langle \text{proof} \rangle\)

**Lemma** \(\text{mon-pl-of-} \alpha nl\): \(\text{mon-pl} (\text{map} (\alpha n fg) w) = \text{mon-loc} fg w \cup \text{mon-env} fg w\)
\(\langle \text{proof} \rangle\)

We now derive specialized introduction lemmas for \(\otimes_{\alpha n fg}\)

**Lemma** \(cil-\alpha n-cons-helper\): \(\text{mon-pl} (\text{map} (\alpha n fg) wb) = \text{mon-ww} fg wb\)
\(\langle \text{proof} \rangle\)

**Lemma** \(cil-\alpha nl-cons-helper\):
\(\text{mon-pl} (\text{map} (\alpha nl fg) wb) = \text{mon-ww} fg (\text{map le-rem-s wb})\)
\(\langle \text{proof} \rangle\)

**Lemma** \(cil-\alpha n-cons1\):
\([w \in wa \otimes_{\alpha n fg} \alpha n fg; \text{fst} (\alpha n fg e) \cap \text{mon-ww} fg wb = \{}]\)
\(\Rightarrow e#w \in e#wa \otimes_{\alpha n fg} \alpha n fg wb\)
\(\langle \text{proof} \rangle\)

**Lemma** \(cil-\alpha n-cons2\):
\([w \in wa \otimes_{\alpha n fg} \alpha n fg; \text{fst} (\alpha n fg e) \cap \text{mon-ww} fg wa = \{}]\)
\(\Rightarrow e#w \in wa \otimes_{\alpha n fg} \alpha n fg e#wb\)
\(\langle \text{proof} \rangle\)

### 8.5.2 Monitors

**Lemma** (in flowgraph) \(ntrs-mon\):
\(\text{assumes} A: ((s,c),e,(s',c')) \in ntrs fg\)
\(\text{shows} \text{mon-s} fg s' = \text{mon-s} fg s \cup \text{fst} (\alpha n fg e)\)
\(\langle \text{proof} \rangle\)

**Corollary** (in flowgraph) \(ntr-called-mon\):
\(\text{assumes} A: ((s,c),e,(s',c')) \in ntrs fg\)
\(\text{shows} \text{fst} (\alpha n fg e) \subseteq \text{mon-s} fg s'\)
\(\langle \text{proof} \rangle\)

**Lemma** (in flowgraph) \(ntr-mon\):
\((c,e,c') \in ntr fg \Rightarrow \text{mon-c} fg c' = \text{mon-c} fg c \cup \text{fst} (\alpha n fg e)\)
\(\langle \text{proof} \rangle\)

**Lemma** (in flowgraph) \(ntrp-mon\):
\(\text{assumes} A: ((s,c),e,(s',c')) \in ntrp fg\)
\(\text{shows} \text{mon-c} fg (\{#s#\} + c) = \text{mon-c} fg (\{#s#\} + c) \cup \text{fst} (\alpha nI fg e)\)
\(\langle \text{proof} \rangle\)
8.5.3 Interleaving theorem

In this section, we show that the consistent interleaving operator describes the intuition behind interleavability of normalized paths. We show: Two paths are simultaneously executable if and only if they are consistently interleaveable and the monitors of the initial configurations are compatible.

The split lemma splits an execution from a context of the form $ca + cb$ into two interleavable executions from $ca$ and $cb$ respectively. While further down we prove this lemma for loc/env-path, which is more general but also more complicated, we start with the proof for paths of the multiset-semantics for illustrating the idea.

**Lemma (in flowgraph) ntr-split:**

\[
\forall ca cb. \ [(ca+cb,w,c') \in \text{trcl (ntr fg)}; \ \text{valid fg (ca+cb)}] \implies \\
\exists ca' cb' wa wb. \ c'=(ca'+cb') \land \\
\text{w} \in (wa \otimes_\alpha f_n fg \text{ wb}) \land \\
\text{mon-c fg ca} \cap (\text{mon-c fg cb} \cup \text{mon-ww fg wb}) = \{} \land \\
\text{mon-c fg cb} \cap (\text{mon-c fg ca} \cup \text{mon-ww fg wa}) = \{} \land \\
(ca,wa,ca') \in \text{trcl (ntr fg)} \land (cb,wb,cb') \in \text{trcl (ntr fg)}
\]

\langle \text{proof} \rangle

The next lemma is a more general version of flowgraph.ntr-split for the semantics with a distinguished local thread. The proof follows exactly the same ideas, but is more complex.

**Lemma (in flowgraph) ntrp-split:**

\[
\forall s c1 c2 s' c'. \ [(s,c1+c2),w,(s',c')] \in \text{trcl (ntrp fg)}; \ \text{valid fg (\{#s\}+c1+c2)}] \implies \\
\exists w1 w2 c1' c2'. \ c'=c1'+c2' \land \\
w \in w1 \otimes_\alpha \text{nl fg (map ENV w2)} \land \\
(s,c1),w1,(s',c1') \in \text{trcl (ntrp fg)} \land \\
(c2,w2,c2') \in \text{trcl (ntrp fg)} \land \\
\text{mon-ww fg (map le-rem-s w1)} \cap \text{mon-c fg c2} = \{} \land \\
\text{mon-ww fg w2} \cap \text{mon-c fg (\{#s\}+c1)} = \{}
\]

\langle \text{proof} \rangle

**Lemma (in flowgraph) ntr-split':**

- **Assumes A:** $(ca+cb,w,c') \in \text{trcl (ntr fg)}$
- **Valid:** $\\text{valid fg (ca+cb)}$
- **Shows:** $\exists ca' cb' wa wb. \ c'=(ca'+cb') \land \\
w \in (wa \otimes_\alpha f_n fg \text{ wb}) \land \\
\text{mon-c fg ca} \cap (\text{mon-c fg cb} \cup \text{mon-ww fg wb}) = \{} \land \\
\text{mon-c fg cb} \cap (\text{mon-c fg ca} \cup \text{mon-ww fg wa}) = \{} \land \\
(ca,wa,ca') \in \text{trcl (ntr fg)} \land \\
(cb,wb,cb') \in \text{trcl (ntr fg)}$

\langle \text{proof} \rangle
The unsplit lemma combines two interleavable executions. For illustration purposes, we first prove the less general version for multiset-configurations. The general version for loc/env-configurations is shown later.

**lemma (in flowgraph)** \(\text{ntr-unsplit} \):  
assumes \( A: w \in wa \otimes_{\text{onl}} fg \ni wb \) and  
\( B: (ca,wa,ca') \in \text{trcl}(\text{ntr} \ fg) \)  
\( (cb,wb,cb') \in \text{trcl}(\text{ntr} \ fg) \)  
\( \text{mon-c} \ fg \ ca \cap (\text{mon-c} \ fg \ cb \cup \text{mon-ww} \ fg \ wb) = \{\} \)  
\( \text{mon-c} \ fg \ cb \cap (\text{mon-c} \ fg \ ca \cup \text{mon-ww} \ fg \ wa) = \{\} \)  
shows \( (ca + cb, w, ca' + cb') \in \text{trcl}(\text{ntr} \ fg) \)  
(proof)

**lemma (in flowgraph)** \(\text{ntrp-unsplit} \):  
assumes \( A: w \in wa \otimes_{\text{onl}} fg \) (map \(\text{ENV} \) \(\ni\) wb) and  
\( B: ((s, ca), wa, (s', ca')) \in \text{trcl}(\text{ntrp} \ fg) \)  
\( (cb, wb, cb') \in \text{trcl}(\text{ntr} \ fg) \)  
\( \text{mon-c} \ fg \ ((\#s\#) + ca) \cap (\text{mon-c} \ fg \ cb \cup \text{mon-ww} \ fg \ wb) = \{\} \)  
\( \text{mon-c} \ fg \ cb \cap (\text{mon-c} \ fg \ ((\#s\#) + ca) \cup \text{mon-ww} \ fg \ wa) = \{\} \)  
shows \( ((s, ca + cb), w, (s', ca' + cb')) \in \text{trcl}(\text{ntrp} \ fg) \)  
(proof)

And finally we get the desired theorem: *Two paths are simultaneously executable if and only if they are consistently interleavable and the monitors of the initial configurations are compatible*. Note that we have to assume a valid starting configuration.

**theorem (in flowgraph)** \(\text{ntr-interleave} \): \(\text{valid} \ fg \ (ca + cb) \implies \)  
\( (ca + cb, w, c') \in \text{trcl}(\text{ntr} \ fg) \iff \)  
\( \exists ca' \text{ } cb' \text{ } wa \text{ wb} . \)  
\( c' = ca + cb' \land \)  
\( w \in (wa \otimes_{\text{onl}} fg \ni wb) \land \)  
\( \text{mon-c} \ fg \ ca \cap (\text{mon-c} \ fg \ cb \cup \text{mon-ww} \ fg \ wb) = \{\} \land \)  
\( \text{mon-c} \ fg \ cb \cap (\text{mon-c} \ fg \ ca \cup \text{mon-ww} \ fg \ wa) = \{\} \land \)  
\( (ca,wa,ca') \in \text{trcl}(\text{ntr} \ fg) \land (cb,wb,cb') \in \text{trcl}(\text{ntr} \ fg) \)  
(proof)

**theorem (in flowgraph)** \(\text{ntrp-interleave} \): \(\text{valid} \ fg \ ((\#s\#) + c1 + c2) \implies \)  
\( ((s,c1 + c2), w, (s', c') \in \text{trcl}(\text{ntrp} \ fg) \iff \)  
\( \exists w1 \text{ } w2 \text{ } c1' \text{ } c2' . \)  
\( w \in w1 \otimes_{\text{onl}} fg \) (map \(\text{ENV} \) \(\ni\) w2) \land \)  
\( c' = c1' + c2' \land \)  
\( ((s,c1), w1, (s', c1') \in \text{trcl}(\text{ntrp} \ fg) \land \)  
\( (c2, w2, c2') \in \text{trcl}(\text{ntrp} \ fg) \land \)  
\( \text{mon-ww} \ fg \) (map \(\text{le-rem-s} \) \(\ni\) w1) \cap \)  
\( \text{mon-c} \ fg \ c2 = \{\} \land \)  
\( \text{mon-ww} \ fg \ w2 \cap \text{mon-c} \ fg \ ((\#s\#) + c1) = \{\} \)  
(proof)
The next is a corollary of flowgraph.ntrp-unsplit, allowing us to convert a path to loc/env semantics by adding a local stack that does not make any steps.

**corollary** (in flowgraph) ntr2ntrp: \[
(c,w,c') \in \text{trcl} (\text{ntr} fg); \\
\text{mon-c fg} (\{\#s\#\}+cl) \cap (\text{mon-c fg} c \cup \text{mon-ww fg} w) = \{\} \\
\implies ((s,cl+c'), \text{map ENV} w,(s,cl+c')) \in \text{trcl} (\text{ntrp} fg)
\]

**8.5.4 Reverse splitting**

This section establishes a theorem that allows us to find the thread in the original configuration that created some distinguished thread in the final configuration.

**lemma** (in flowgraph) ntr-reverse-split: \![w s' ce']. \[
(c,w,\{\#s\#\}+ce') \in \text{trcl} (\text{ntr} fg); \\
\text{valid fg} c \implies \exists s ce w1 w2 ce1' ce2'. \\
c = \{\#s\#\} + ce \land \\
ce' = ce1' + ce2' \land \\
w \in w1 \odot_{\text{on} fg} w2 \land \\
\text{mon-s fg} s \cap (\text{mon-c fg} ce \cup \text{mon-ww fg} w2) = \{\} \land \\
\text{mon-c fg} ce \cap (\text{mon-s fg} s \cup \text{mon-ww fg} w1) = \{\} \land \\
(\{\#s\#\}, w1, \{\#s'\#\} + ce1') \in \text{trcl} (\text{ntr} fg) \land \\
(ce, w2, ce2') \in \text{trcl} (\text{ntr} fg)
\]

— The proof works by induction on the initial configuration. Note that configurations consist of finitely many threads only
— FIXME: An induction over the size (rather then over the adding of some fixed element) may lead to a smoother proof here

**end**

**9 Constraint Systems**

**theory** ConstraintSystems

**imports** Main AcquisitionHistory Normalization

**begin**

In this section we develop a constraint-system-based characterization of our analysis.

Constraint systems are widely used in static program analysis. There least solution describes the desired analysis information. In its generic form, a constraint system \( R \) is a set of inequations over a complete lattice \((L, \sqsubseteq)\) and a set of variables \( V \). An inequation has the form \( R[v] \sqsupseteq \text{rhs} \), where \( R[v] \in V \) and \( \text{rhs} \) is a monotonic function over the variables. Note that
for program analysis, there is usually one variable per control point. The variables are then named $R[v]$, where $v$ is a control point. By standard fixed-point theory, those constraint systems have a least solution. Outside the constraint system definition $R[v]$ usually refers to a component of that least solution.

Usually a constraint system is generated from the program. For example, a constraint generation pattern could be the following:

$$
\begin{array}{l}
\text{for } (u, \text{Call } q, v) \in E: \\
\quad S^k[v] \supseteq \{ (\text{mon}(q) \cup M \cup M', \hat{P}) | (M, P) \in S^k[u] \land (M', P') \in S^k[r_q] \\
\quad \land \hat{P} \leq P \cup P' \land |\hat{P}| \leq 2 \} \\
\end{array}
$$

For some parameter $k$ and a flowgraph with nodes $N$ and edges $E$, this generates a constraint system over the variables $\{ S^k[v] \mid v \in N \}$. One constraint is generated for each call edge. While we use a powerset lattice here, we can in general use any complete lattice. However, all the constraint systems needed for our conflict analysis are defined over powerset lattices $(P('a), \subseteq)$ for some type 'a. This admits a convenient formalization in Isabelle/HOL using inductively defined sets. We inductively define a relation between variables\(^3\) and the elements of their values in the least solution, i.e. the set $\{(v, x) \mid x \in R[v]\}$. For example, the constraint generator pattern from above would become the following introduction rule in the inductive definition of the set $S-cs \ fg \ k$:

$$
\begin{array}{l}
\left[ \left( (u, \text{Call } q, v) \in \text{edges } fg; \ (u, M, P) \in S-cs \ fg \ k; \ (\text{return } fg \ q, Ms, Ps) \in S-cs \ fg \ k; \ P' \leq \#P + Ps; \ \text{size } P' \leq k \right) \right] \\
\implies (v, \text{mon } fg \ q \cup M \cup Ms, P') \in S-cs \ fg \ k
\end{array}
$$

The main advantage of this approach is that one gets a concise formalization by using Isabelle’s standard machinery, the main disadvantage is that this approach only works for powerset lattices ordered by $\subseteq$.

### 9.1 Same-level paths

#### 9.1.1 Definition

We define a constraint system that collects abstract information about same-level paths. In particular, we collect the set of used monitors and all multi-subsets of spawned threads that are not bigger than $k$ elements, where $k$ is a parameter that can be freely chosen.

An element $(u, M, P) \in S-cs \ fg \ k$ means that there is a same-level path from the entry node of the procedure of $u$ to $u$, that uses the monitors $M$ and spawns at least the threads in $P$.

---

\(^3\)Variables are identified by control nodes here
The intuition underlying this constraint system is the following: The $S$-init-constraint describes that the procedures entry node can be reached with the empty path, that has no monitors and spawns no procedures. The $S$-base-constraint describes that executing a base edge does not use monitors or spawn threads, so each path reaching the start node of the base edge also induces a path reaching the end node of the base edge with the same set of monitors and the same set of spawned threads. The $S$-call-constraint models the effect of a procedure call. If there is a path to the start node of a call edge and a same-level path through the procedure, this also induces a path to the end node of the call edge. This path uses the monitors of both path and spawns the threads that are spawned on both paths. Since we only record a limited subset of the spawned threads, we have to choose which of the threads are recorded. The $S$-spawn-constraint models the effect of a spawn edge. A path to the start node of the spawn edge induces a path to the end node that uses the same set of monitors and spawns the threads of the initial path plus the one spawned by the spawn edge. We again have to choose which of these threads are recorded.

9.1.2 Soundness and Precision

Soundness of the constraint system $S$-cs means, that every same-level path has a corresponding entry in the constraint system.

As usual the soundness proof works by induction over the length of execution paths. The base case (empty path) trivially follows from the $S$-init constraint. In the inductive case, we consider the edge that induces the last step of the path; for a return step, this is the corresponding call edge (cf. Lemma flowgraph.trss-find-call'). With the induction hypothesis, we get the soundness for the (shorter) prefix of the path, and depending on the last step we can choose a constraint that implies soundness for the whole path.

**lemma** (in flowgraph) $S$-sound: $!p v c' P$. 
Precision means that all entries appearing in the smallest solution of the constraint system are justified by some path in the operational characterization. For proving precision, one usually shows that a family of sets derived as an abstraction from the operational characterization solves all constraints.

In our formalization of constraint systems as inductive sets this amounts to constructing for each constraint a justifying path for the entries described on the conclusion side of the implication — under the assumption that corresponding paths exists for the entries mentioned in the antecedent.

**lemma (in flowgraph) S-precise:** \((v,M,P) \in S\)-cs \(fg k\) \(\Rightarrow \exists p c' w. (([entry \ fg \ p],[\#]),w,([v],c')) \in \text{trcl} (trss \ fg) \land \)

\[
\begin{align*}
&\text{size } P \leq k \land \ \\
&\lambda p. [entry \ fg \ p] \ '## P \leq c' \land \\
&M=\text{mon-w} \ fg \ w
\end{align*}
\]

**proof**

**theorem (in flowgraph) S-sound-precise:**

\((v,M,P) \in S\)-cs \(fg k\) \(\iff\)

\[
\begin{align*}
&\exists p c' w. (([entry \ fg \ p],[\#]),w,([v],c')) \in \text{trcl} (trss \ fg) \land \ \\
&\text{size } P \leq k \land (\lambda p. [entry \ fg \ p]) \ '## P \leq c' \land \\
&M=\text{mon-w} \ fg \ w
\end{align*}
\]

**proof**

Next, we present specialized soundness and precision lemmas, that reason over a macrostep \((ntrp \ fg)\) rather than a same-level path \((\text{trcl} \ (trss \ fg))\). They are tailored for the use in the soundness and precision proofs of the other constraint systems.

**lemma (in flowgraph) S-sound-ntrp:**

**assumes A:** \([(u),[\#]),(sh,ch)]\) \(\in ntrp \ fg\) \(\land\)

**case** \(!!! p u v w. [ u, Spawn p,v] \in edges \ fg \land \\

sh=[v,u']; \\

\text{proc-of f g v} = p; \\

\text{mon-c f g ch} = \{\};

!!s. s:\# ch \Rightarrow \exists p u v. s=[entry \ fg \ p] \land \\

\text{initialproc f g p}; \\

!!P. (\lambda p. [entry \ fg \ p]) \ '## P \leq ch \Rightarrow \\

(v,mon-w \ fg \ w,P) \in S\)-cs \(fg k\) \((\text{size } P)\)

\]

**shows Q**

**proof**

**lemma (in flowgraph) S-precise-ntrp:**
assumes $ENTRY: (v,M,P) \in S-cs fg k$ and

$P: \text{proc-of } fg \ v = p$ and

$EDGE: (u,\text{Call } p,u^\prime) \in \text{edges } fg$

shows $\exists w ch.
\left( ([u],\{\#\}),\text{LOC } (\text{LCall } p\#w),([v,u^\prime],ch) \right) \in \text{ntrp } fg \land
\text{size } P \leq k \land
M = \text{mon-w } fg \ w \land
\text{mon-n } fg \ v = \text{mon-fg } p \land
(\lambda p. [\text{entry } fg \ p]) \ # P \leq ch \land
\text{mon-c } fg \ ch = \{\}
\right)$

9.2 Single reaching path

In this section we define a constraint system that collects abstract information of paths reaching a control node at $U$. The path starts with a single initial thread. The collected information are the monitors used by the steps of the initial thread, the monitors used by steps of other threads and the acquisition history of the path. To distinguish the steps of the initial thread from steps of other threads, we use the loc/env-semantics (cf. Section 5.4).

9.2.1 Constraint system

An element $(u, M_l, M_e, h) \in RU-cs fg U$ corresponds to a path from $\{\#u\#\}$ to some configuration at $U$, that uses monitors from $M_l$ in the steps of the initial thread, monitors from $M_e$ in the steps of other threads and has acquisition history $h$.

Here, the correspondence between paths and entries included into the inductively defined set is not perfect but strong enough for our purposes: While each constraint system entry corresponds to a path, not each path corresponds to a constraint system entry. But for each path reaching a configuration at $U$, we find an entry with less or equal monitors and an acquisition history less or equal to the acquisition history of the path.

inductive-set

$RU-cs : (n',p,'ba','m','more) \text{flowgraph-rec-scheme } \Rightarrow \ n \ set \Rightarrow
\left( (n' \times 'm set \times 'm set \times ('m \Rightarrow 'm set)) \ set
\right)$

for $fg U$

where

$RU\text{-init: } u \in U \implies (u,\{\},\lambda x.\{\}) \in RU-cs fg U$

$RU\text{-call: } \left( \begin{array}{c}
(u,\text{Call } p,u^\prime) \in \text{edges } fg; \ \text{proc-of } fg \ v = p; \ (v,M,P) \in S-cs fg 0; \\
(v,M_l,M_e,h) \in RU-cs fg U; \ \text{mon-n } fg \ u \cap M_e = \{\}
\end{array} \right) \implies (u,\ \text{mon-fg } p \cup M \cup M_l, M_e, \text{ah-update } h (\text{mon-fg } p,M) (M_l \cup M_e)) \in RU-cs fg U$

$RU\text{-spawn: } \left( \begin{array}{c}
(u,\text{Call } p,u^\prime) \in \text{edges } fg; \ \text{proc-of } fg \ v = p; \ (v,M,P) \in S-cs fg 1; \\
q\#P; \ (\text{entry } fg \ q,M_l,M_e,h) \in RU-cs fg U; \\
(\text{mon-n } fg \ u \cup \text{mon-fg } p) \cap (M_l \cup M_e) = \{\}
\end{array} \right)$
The constraint system works by tracking only a single thread. Initially, there is just one thread, and from this thread we reach a configuration at $U$. After a macrostep, we have the transformed initial thread and some spawned threads. The key idea is, that the actual node $U$ is reached by just one of these threads. The steps of the other threads are useless for reaching $U$. Because of the nice properties of normalized paths, we can simply prune those steps from the path.

The $RU$-init-constraint reflects that we can reach a control node from itself with the empty path. The $RU$-call-constraint describes the case that $U$ is reached from the initial thread, and the $RU$-spawn-constraint describes the case that $U$ is reached from one of the spawned threads. In the two latter cases, we have to check whether prepending the macrostep to the reaching path is allowed or not due to monitor restrictions. In the call case, the procedure of the initial node must not own monitors that are used in the environment steps of the appended reaching path ($\text{mon-n fg u} \cap \text{Me} = \{\}$). As we only test disjointness with the set of monitors used by the environment, reentrant monitors can be handled. In the spawn case, we have to check disjointness with both, the monitors of local and environment steps of the reaching path from the spawned thread, because from the perspective of the initial thread, all these steps are environment steps ($((\text{mon-n fg u} \cup \text{mon fg p}) \cap (\text{Ml} \cup \text{Me})) = \{\}$). Note that in the call case, we do not need to explicitly check that the monitors used by the environment are disjoint from the monitors acquired by the called procedure because this already follows from the existence of a reaching path, as the starting point of this path already holds all these monitors.

However, in the spawn case, we have to check for both the monitors of the start node and of the called procedure to be compatible with the already known reaching path from the entry node of the spawned thread.

### 9.2.2 Soundness and precision

The following lemma intuitively states: If we can reach a configuration that is at $U$ from some start configuration, then there is a single thread in the start configuration that can reach a configuration at $U$ with a subword of the original path.

The proof follows from Lemma `flowgraph.ntr-reverse-split` rather directly.

**Lemma (in `flowgraph`) ntr-reverse-split-atU:**

```
assumes V: valid fg c and
A: atU U c' and
B: (c,w,c')\in trcl (ntr fg)
sows \exists s w' c1'.
```

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The next lemma shows the soundness of the RU constraint system.

The proof works by induction over the length of the reaching path. For the empty path, the proposition follows by the RU-init-constraint. For a non-empty path, we consider the first step. It has transformed the initial thread and may have spawned some other threads. From the resulting configuration, $U$ is reached. Due to flowgraph.ntr-split we get two interleavable paths from the rest of the original path, one from the transformed initial thread and one from the spawned threads. We then distinguish two cases: if the first path reaches $U$, the proposition follows by the induction hypothesis and the RU-call constraint.

Otherwise, we use flowgraph.ntr-reverse-split-atU to identify the thread that actually reaches $U$ among all the spawned threads. Then we apply the induction hypothesis to the path of that thread and prepend the first step using the RU-spawn-constraint.

The main complexity of the proof script below results from fiddling with the monitors and converting between the multiset-and loc/env-semantics. Also the arguments to show that the acquisition histories are sound approximations require some space.

**Lemma (in flowgraph) RU-sound:**

\[
\begin{align*}
& s : \# c \land w' \leq w \land c_1' \leq c' \land \\
& atU U c_1' \land (\{\# s'\}, w', c_1') \in \text{trcl (ntr f)}
\end{align*}
\]

(proof)

Now we prove a statement about the precision of the least solution. As in the precision proof of the S-cs constraint system, we construct a path for the entry on the conclusion side of each constraint, assuming that there already exists paths for the entries mentioned in the antecedent.

We show that each entry in the least solution corresponds exactly to some executable path, and is not just an under-approximation of a path; while for the soundness direction, we could only show that every executable path is under-approximated. The reason for this is that in effect, the constraint system prunes the steps of threads that are not needed to reach the control point. However, each pruned path is executable.

**Lemma (in flowgraph) RU-precise:**

\[
\begin{align*}
& (u, s', c') \land (([u], \{\#\}), w, (s', c')) \in \text{trcl (ntr f)}; atU U \{\# s'\} + c'
\end{align*}
\]

(proof)
9.3 Simultaneously reaching path

In this section, we define a constraint system that collects abstract information for paths starting at a single control node and reaching two program points simultaneously, one from a set $U$ and one from a set $V$.

9.3.1 Constraint system

An element $(u, Ml, Me) \in RUV\text{-}cs fg U V$ means, that there is a path from $\{\#u\#\}$ to some configuration that is simultaneously at $U$ and at $V$. That path uses monitors from $Ml$ in the first thread and monitors from $Me$ in the other threads.

\begin{align*}
\text{inductive-set} & \\
& RUV\text{-}cs :: (\text{\textquotesingle\textquotesingle}n', \text{\textquotesingle\textquotesingle}p', \text{\textquotesingle\textquotesingle}ba', \text{\textquotesingle\textquotesingle}m', \text{\textquotesingle\textquotesingle}more) \text{ flowgraph-rec-scheme} \Rightarrow \\
& \quad \text{\textquotesingle\textquotesingle}n\text{ set} \Rightarrow \text{\textquotesingle\textquotesingle}n\text{ set} \Rightarrow (\text{\textquotesingle\textquotesingle}n \times \text{\textquotesingle\textquotesingle}m\text{ set} \times \text{\textquotesingle\textquotesingle}m\text{ set}) \text{ set}
\end{align*}

\textbf{for } fg U V

\textbf{where}

\begin{align*}
& RUV\text{-}call: \\
& \quad \exists (u, \text{\textquotesingle\textquotesingle}Call p, u') \in \text{edges } fg; \text{ proc-of } fg v = p; (v, Ml, Me) \in S\text{-}cs fg 0; \\
& \quad (v, Ml, Me) \in RUV\text{-}cs fg U V; \text{ mon-}n\ fg u \cap Me = \{\} \quad \Rightarrow \quad (u, \text{mon-}n\ fg p \cup M \cup Ml, Me) \in RUV\text{-}cs fg U V
\end{align*}

\begin{align*}
& \text{\textbf{| RUV\text{-}spawn:}} \\
& \quad \exists (u, \text{\textquotesingle\textquotesingle}Call p, u') \in \text{edges } fg; \text{ proc-of } fg v = p; (v, Ml, Me) \in S\text{-}cs fg 1; q : \# P; \\
& \quad (entry\ fg\ q, Ml, Me) \in RUV\text{-}cs fg U V; \\
& \quad (\text{mon-}n\ fg u \cup \text{mon-}n\ fg p) \cap (Ml \cup Me) = \{\} \quad \Rightarrow \quad (u, \text{mon-}n\ fg p \cup M \cup Ml, Me) \in RUV\text{-}cs fg U V
\end{align*}

\begin{align*}
& \text{\textbf{| RUV\text{-}split-le:}} \\
& \quad \exists (u, \text{\textquotesingle\textquotesingle}Call p, u') \in \text{edges } fg; \text{ proc-of } fg v = p; (v, Ml, Me, h) \in RU\text{-}cs fg U; (entry\ fg\ q, Ml', Me', h') \in RU\text{-}cs fg V; \\
& \quad (\text{mon-}n\ fg u \cup \text{mon-}n\ fg p) \cap (Me \cup Ml' \cup Me') = \{\}; h \ [s]\ h' \quad \Rightarrow \quad (u, \text{mon-}n\ fg p \cup M \cup Ml, Me \cup Ml' \cup Me') \in RUV\text{-}cs fg U V
\end{align*}

\begin{align*}
& \text{\textbf{| RUV\text{-}split-e:}} \\
& \quad \exists (u, \text{\textquotesingle\textquotesingle}Call p, u') \in \text{edges } fg; \text{ proc-of } fg v = p; (v, Ml, Me) \in S\text{-}cs fg 2; \\
& \quad (\#q\#) + (\#q'\#) \leq P; \\
& \quad (entry\ fg\ q, Ml, Me, h) \in RU\text{-}cs fg U; (entry\ fg\ q', Ml', Me', h') \in RU\text{-}cs fg V; \\
& \quad (\text{mon-}n\ fg u \cup \text{mon-}n\ fg p) \cap (Me \cup Ml \cup Ml' \cup Me') = \{\}; h \ [s]\ h' \quad \Rightarrow \quad (u, \text{mon-}n\ fg p \cup M \cup Ml, Me \cup Ml' \cup Me') \in RUV\text{-}cs fg U V
\end{align*}
The idea underlying this constraint system is similar to the RU-cs-constraint system for reaching a single node set. Initially, we just track one thread. After a macrostep, we have a configuration consisting of the transformed initial thread and the spawned threads. From this configuration, we reach two nodes simultaneously, one in $U$ and one in $V$. Each of these nodes is reached by just a single thread. The constraint system contains one constraint for each case how these threads are related to the initial and the spawned threads:

**RUVCall** Both, $U$ and $V$ are reached from the initial thread.

**RUVPawn** Both, $U$ and $V$ are reached from a single spawned thread.

**RUVSplitLe** $U$ is reached from the initial thread, $V$ is reached from a spawned thread.

**RUVSplitEl** $V$ is reached from the initial thread, $U$ is reached from a spawned thread.

**RUVSplitEe** Both, $U$ and $V$ are reached from different spawned threads.

In the latter three cases, we have to analyze the interleaving of two paths each reaching a single control node. This is done via the acquisition history information that we collected in the RU-cs-constraint system.

Note that we do not need an initializing constraint for the empty path, as a single configuration cannot simultaneously be at two control nodes.

### 9.3.2 Soundness and precision

**Lemma (in flowgraph)** RUVC-sound: $!!u \ s' \ c'$.

\[
\begin{align*}
\Rightarrow & \exists Ml \ Me. \\
& (u, Ml, Me) \in RUV-cs \ fg \ U \ V \land \\
& Ml \subseteq \text{mon-loc fg w} \\
& Me \subseteq \text{mon-env fg w}
\end{align*}
\]

— The soundness proof is done by induction over the length of the reaching path

**Lemma (in flowgraph)** RUVC-precise: $(u, Ml, Me) \in RUV-cs \ fg \ U \ V$

\[
\begin{align*}
\Rightarrow & \exists w \ s' \ c'. \\
& (([u],\{\#\}),w,(s',c')) \in \text{trcl (ntrp \ fg)} \land \\
& atUV \ U \ V \ ((\#s'\#)+c') \land \\
& \text{mon-loc fg w} = Ml \land \\
& \text{mon-env fg w} = Me
\end{align*}
\]

**Proof**

end
10 Main Result

theory MainResult
imports ConstraintSystems
begin

At this point everything is available to prove the main result of this project: 
*The constraint system RUV-cs precisely characterizes simultaneously reachable control nodes w.r.t. to our semantic reference point.*

The „trusted base“ of this proof, that are all definitions a reader that trusts the Isabelle prover must additionally trust, is the following:

- The flowgraph and the assumptions made on it in the flowgraph- and eflowgraph-locales. Note that we show in Section 6.4 that there is at least one non-trivial model of eflowgraph.
- The reference point semantics (refpoint) and the transitive closure operator (trcl).
- The definition of atUV.
- All dependencies of the above definitions in the Isabelle standard libraries.

*theorem* (in eflowgraph) RUV-is-sim-reach:

\[ \exists w c'. (\#[entry fg (main fg)]#, w, c') \in \text{trcl (refpoint fg)} \land \text{atUV U V c'} \]

\[ \iff \exists Ml Me. (entry fg (main fg), Ml, Me) \in \text{RUV-cs fg U V} \]

— The proof uses the soundness and precision theorems wrt. to normalized paths (flowgraph.RUV-sound, flowgraph.RUV-precise) as well as the normalization result, i.e. that every reachable configuration is also reachable using a normalized path (eflowgraph.normalize) and, vice versa, that every normalized path is also a usual path (ntr-is-tr). Finally the conversion between our working semantics and the semantic reference point is exploited (flowgraph.refpoint-eq).

⟨proof⟩

end

11 Conclusion

We have formalized a flowgraph-based model for programs with recursive procedure calls, dynamic thread creation and reentrant monitors and its operational semantics. Based on the operational semantics, we defined a conflict as being able to simultaneously reach two control points from two given sets \(U\) and \(V\) when starting at the initial program configuration, just consisting of a single thread at the entry point of the main procedure. We then formalized a constraint-system-based analysis for conflicts and proved
it sound and precise w.r.t. the operational definition of a conflict. The main idea of the analysis was to restrict the possible schedules of a program. On the one hand, this restriction enabled the constraint system based analysis, on the other hand it did not change the set of reachable configurations (and thus the set of conflicts).

We characterized the constraint systems as inductive sets. While we did not derive an executable algorithm explicitly, the steps from the inductive sets characterization to an algorithm follow the path common in program analysis and pose no particular difficulty. The algorithm would have to construct a constraint system (system of inequalities over a finite height lattice) from a given program corresponding to the inductively defined sets studied here and then determine its least solution, e.g. by a worklist algorithm. In order to make the algorithm executable, we would have to introduce finiteness assumptions for our programs. The derivation of executable algorithms is currently in preparation.

A formal analysis of the algorithmic complexity of the problem will be presented elsewhere. Here we only present some results: Already the problem of deciding the reachability of a single control node is NP-hard, as can be shown by a simple reduction from SAT. On the other hand, we can decide simultaneous reachability in nondeterministic polynomial time in the program size, where the number of random bits depends on the possible nesting depth of the monitors. This can be shown by analyzing the constraint systems.

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References


