Proof-Carrying-Code

Applying formal methods in a distributed world

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Motivation

Resource-bounded computation is one specific instance of PCC.

Safety policy: resource consumption is lower than a given bound.

Resources can be (heap) space, time, or size of parameters to system calls.

Strong demand for such guarantees for example in embedded systems.

Mobile Resource Guarantees

Objective:
Development of an infrastructure to endow mobile code with independently verifiable certificates describing resource behaviour.

Approach:
Proof-carrying code for **resource-related properties**, where proofs are generated from typing derivations in a **resource-aware type system**.

Motivation

Restrict the execution of mobile code to those adhering to a certain resource policy.

Application Scenarios:

- A user of a **handheld device** might want to know that a downloaded application will definitely run within the limited amount of memory available.
- A provider of **computational power in a Grid infrastructure** may only be willing to offer this service upon receiving dependable guarantees about the required resource consumption.

Proof-Carrying-Code with High-Level-Logics

Our approach to PCC: Combine high-level type-systems with program logics and build a **hierarchy of logics** to construct a logic tailored to reason about resources.

Everything is **formalised in a theorem prover**.

Classic vs Foundational PCC: best of both worlds

- **Simple reasoning**, using specialised logics;
- **Strong foundations**, by encoding the logics in a theorem prover

Proof-Carrying-Code with High-Level-Logics

High-Level Type System	$G \vdash_{\mathsf{H}} t : \tau$
Specialised Logic	$\rhd \ulcorner t \urcorner : D(G, \tau)$
Termination Logic	$\vdash_{\mathsf{T}} \{P\}\ \mathsf{e}\ \downarrow$
Program Logic	$\Gamma\ \rhd\ \mathsf{e} : A$
Operational Semantics	$E \vdash h, \mathsf{e} \Downarrow (h', v, p)$

Motivating Example of this Hierarchical Approach

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Directly on the program logic

$$\rhd\ f(x)\ :\ \lambda E\ h\ h'\ v\ .\ h \models_{list} E\langle x\rangle\ \longrightarrow\ h' \models_{list} v$$

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NOT: reasoning on this level generates huge side-conditions.

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Instead, define a higher-level logic $\vdash_H$ that abstracts over the details of datatype representation, and that has the property

$$G \vdash_H t : \tau \implies \rhd \ulcorner t \urcorner : D(\Gamma, \tau)$$

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We specialise the form of assertions like this

$$\begin{array}{rcl} D(\{x : list, y : list\}, list) & \equiv & \\ \lambda E\ h\ h'\ v. & & h \models_{list} E\langle x\rangle \ \wedge\ h \models_{list} E\langle y\rangle \ \longrightarrow \\ & & h' \models_{list} E\langle x\rangle \ \wedge\ h' \models_{list} E\langle y\rangle \ \wedge\ h' \models_{list} v \end{array}$$

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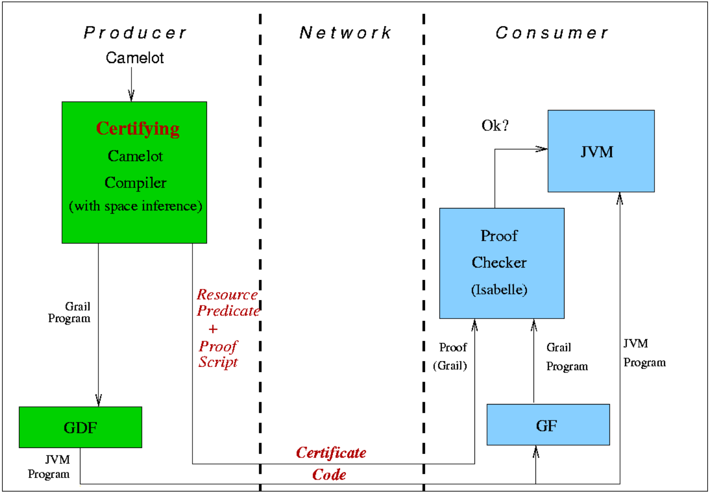
Now we can formulate rules, that match translations from the high-level language:

$$\frac{\rhd \ulcorner t_1 \urcorner : D(\Gamma, \tau) \quad \rhd \ulcorner t_2 \urcorner : D(\Gamma, \tau\ list)}{\rhd \ulcorner cons(t_1, t_2) \urcorner : D(\Gamma, \tau\ list)}$$

A Proof-Carrying-Code Infrastructure for MRG



Camelot

- Strict, first-order functional language with CAML-like syntax and object-oriented extensions
- Compiled to subset of JVM (Java Virtual Machine) bytecode (Grail)
- Memory model: 2 level heap
- Security: Static analyses to prevent deallocation of live cells in Level-1 Heap: linear typing (folklore + Hofmann), readonly typing (Aspinall, Hofmann, Konencny), layered sharing analysis (Konencny).
- Resource bounds: Static analysis to infer linear upper bounds on heap consumption (Hofmann, Jost).

Example: Insertion Sort

Camelot program:

```
let ins a l = match l with
                 Nil -> Cons(a,Nil)
               | Cons(x,t)@_ -> if a < x then Cons(a,Cons(x,t))
                                         else Cons(x, ins a t)

let sort l =  match l with  Nil -> Nil
                          | Cons(a,t)@_ -> ins a (sort t)
```

In-place Operations via a Diamond Type

Using operators, such as `Cons`, amounts to heap allocation.

Additionally, Camelot provides extensions to do **in-place operations** over arbitrary data structures via a so called **diamond type** $\diamond$ with **d** $\in \diamond$:

```
match l with Nil@d => e1
            | Cons (h,t)@d => ... Cons (x,t)@d ...
```

The memory occupied by the cons cell can be **re-used** via the diamond `d`.
Note:

- $\diamond$ is an abstract data-type
- structured use of diamonds in branches of pattern matches

How does this fit with referential transparency?

Using a diamond type, we can introduce side effects:

```
type ilist = Nil | Cons of int*ilist
let insert1 x l =
    match l with Nil -> Cons (x, l)
               | Cons(h,t)@d ->
                   if x <= h then Cons(x, Cons(h,t)@d)
                             else Cons(h, insert1 x t)@d

let sort l =  match l with Nil -> Nil
                         | Cons(h,t) -> insert1 h (sort t)
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let sort l =  match l with Nil -> Nil
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```

Now, what's the result of

```
let start args = let l = [4,5,6,7] in
                 let l1 = insert1 6 l in
                 print_list l
```

Linearity saves the day

We can characterise the class of programs for which referential transparency is retained.

Theorem

A **linearly typed** *Camelot program computes the function specified by its purely functional semantics (Hofmann, 2000).*

Beyond Linearity

But: linearity is too restrictive in many cases; we also want to use diamonds in programs where **only the last access to the data structure is destructive**.

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More expressive type systems to express such access patterns are **readonly types** (Aspinall, Hofmann, Konecny, 2001) and types with **layered sharing** (Konecny 2003).

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As with pointers, diamonds can be a powerful gun to shoot yourself in the foot. We need a **powerful type system** to prevent this, and want a **static analysis** to predict resource consumption.

Space Inference

Goal: Infer a linear upper bound on heap consumption.

Given Camelot program containing a function

```
start : string list -> unit
```

find **linear function** s such that `start`(l) will not call `new()` (only `make()`) when evaluated in a heap h where

- **the freelist has length not less than** $s(n)$
- l points in h to a linear list of some length n
- the freelist which forms a part of h is well-formed
- the freelist does not overlap with l

Composing `start` with *runtime environment* that binds input to, e.g., `stdin` yields a standalone program that runs within predictable heap space.

Extended (LFD) Types

Idea: Weights are attached to constructors in an extended type-system.

```
ins          : 1, int -> list(...<0>) -> list(...<0>), 0
```

says that the call `ins x xs` requires 1 heap-cell plus 0 heap cells for each `Cons` cell of the list `xs`.

```
sort         : 0, list(...<0>) -> list(...<0>), 0
```

`sort` does not consume any heap space.

```
start        : 0, list(...<1>) -> unit, 0;
```

gives rise to the desired linear bounding function $s(n) = n$.

High-level Type System: Function Call

A, B, C are types, $k, k', n, n' \in \mathbb{Q}^+$, $\mathtt{f}$ is a Camelot function and $\mathtt{x_1}, \ldots, \mathtt{x_p}$ are variables, Σ is a table mapping function names to types.

$$\frac{\begin{array}{c}\Sigma(\mathtt{f}) = (A_1, \ldots, A_p, k) \longrightarrow (C, k') \\ n \geq k \qquad n - k + k' \geq n'\end{array}}{\Gamma, \mathtt{x_1} : A_1, \ldots, \mathtt{x_p} : A_p, n \vdash \mathtt{f}(\mathtt{x_1}, \ldots, \mathtt{x_p}) : C, n'} \quad (\textsc{Fun})$$

Grail

Grail is an abstraction over virtual machine languages such as the JVM.

$$
\begin{array}{rcl}
e \in \mathit{expr} & ::= & \texttt{null} \mid \texttt{int}\ i \mid \texttt{var}\ x \mid \texttt{prim}\ p\ x\ x \mid \texttt{new}\ c\ [t_1 := x_1, \ldots, t_n := x_n] \mid \\
& & x.t \mid x.t{:=}x \mid c \diamond t \mid c \diamond t{:=}x \mid \texttt{let}\ x = e\ \texttt{in}\ e \mid e\ ;\ e \mid \\
& & \texttt{if}\ x\ \texttt{then}\ e\ \texttt{else}\ e \mid \texttt{call}\ f \mid x \cdot m(\bar{a}) \mid c \diamond m(\bar{a}) \\
a \in \mathit{args} & ::= & \texttt{var}\ x \mid \mathsf{null} \mid i
\end{array}
$$

Example: Insertion sort

Grail code:

```
method static public List ins (int a, List l) = ...Make(..,..,..
method static public List sort (List l) =
  let fun f(List l) =
    if l = null then null
                else let val h = l.HD
                         val t = l.TL
                         val () = D.free (l)
                         val l = List.sort (t)
                     in List.ins (h, l) end
  in f(l) end
```

This is a 1-to-1 translation of JVM code

Judgement of the Operational Semantics

Modelling resources: Resources are an extra component in operational and axiomatic semantics ("resource record").

$$\mathbf{p} \in RRec = (\!| clock : nat, callcount : nat, invokedepth : nat, maxstack : nat |\!)$$

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A judgement in the functional operational semantics

$$E \vdash h, e \Downarrow_n (h', v, p)$$

is to be read as "starting with a heap h and a variable enviroment E, the Grail code e evaluates in n steps to the value v, yielding the heap h' as result and consuming p resources."

Operational Semantics: Let- and Call-rules

$$\frac{E \vdash h, e_1 \Downarrow_n (h_1, w, p) \quad w \neq \bot \quad E\langle x := w\rangle \vdash h_1, e_2 \Downarrow_m (h_2, v, q)}{E \vdash h, \texttt{let}\ x = e_1\ \texttt{in}\ e_2 \Downarrow_{max(n,m)+1} (h_2, v, \mathbf{p_1} \smile \mathbf{p_2})}\ (\text{LET})$$

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$$\frac{E \vdash h, body_f \Downarrow_n (h_1, v, p)}{E \vdash h, \texttt{call}\ f \Downarrow_{n+1} (h_1, v, \langle \mathbf{1\ 1\ 0\ 0} \rangle \oplus \mathbf{p_1})} \quad (\text{CALL})$$

A Program Logic for Grail

VDM-style logic with judgements of the form $\Gamma \ \rhd \ e : A$, meaning

"in context Γ expression e fulfills the assertion A"

Type of assertions (**shallow embedding**):

$$\mathcal{A} \equiv \mathcal{E} \rightarrow \mathcal{H} \rightarrow \mathcal{H} \rightarrow \mathcal{V} \rightarrow \mathcal{R} \rightarrow \mathcal{B}$$

No syntactic separation into pre- and postconditions.

Semantic validity $\models e : A$ means

"whenever $E \vdash h, e \Downarrow (h', v, p)$ then $A\ E\ h\ h'\ v\ p$ holds"

Note: Covers partial correctness; termination orthogonal.

A Program Logic for Grail

Simplified rule for parameterless function call:

$$\frac{\Gamma, (\texttt{Call f} : A) \;\rhd\; \texttt{e} \;:\; A^{+}}{\Gamma \;\rhd\; \texttt{Call f} \;: A} \qquad (\textsc{CallRec})$$

where $\texttt{e}$ is the body of the function $\texttt{f}$ and

$$A^{+} \equiv \lambda\, E\, h\, h'\, v\, p.\, A(E, h, h', v, p^{+})$$

where p^{+} is the updated cost component.
Note:

- Context Γ: collects hypothetical judgements for recursion
- Meta-logical guarantees: soundness, relative completeness

Program Logic Rules

$$\frac{\Gamma \rhd e_1 : P \quad \Gamma \rhd e_2 : Q}{\begin{array}{ll}\Gamma \rhd \texttt{let}\ x = e_1\ \texttt{in}\ e_2 : \lambda\, E\, h\, h'\, v\, p.\, \exists\ p_1\ p_2\ h_1\ w. & P\ E\ h\ h_1\ w\ p_1\ \wedge\ w \neq \bot\ \wedge \\ & Q\ (E\langle x := w\rangle)\ h_1\ h'\ v\ p_2)\ \wedge \\ & p = \mathbf{p_1} \smile \mathbf{p_2}\end{array}} \quad (\textsc{vlet})$$

Program Logic Rules

$$\frac{\Gamma \rhd e_1 : P \quad \Gamma \rhd e_2 : Q}{\begin{array}{l}\Gamma \rhd \texttt{let } x = e_1 \texttt{ in } e_2 : \lambda E\, h\, h'\, v\, p.\, \exists\ p_1\ p_2\ h_1\ w.\ \ P\ E\ h\ h_1\ w\ p_1\ \wedge\ w \neq \bot \wedge \\ \qquad\qquad\qquad\qquad\qquad\qquad\qquad\qquad\quad Q\ (E\langle x := w\rangle)\ h_1\ h'\ v\ p_2) \wedge \\ \qquad\qquad\qquad\qquad\qquad\qquad\qquad\qquad\quad p = \mathbf{p_1} \smile \mathbf{p_2}\end{array}} \quad (\text{VLET})$$

$$\frac{\Gamma \cup \{(\texttt{call } f, P)\} \rhd body_f : \lambda E\, h\, h'\, v\, p.\, P\ E\ h\ h'\ v\ \langle \mathbf{1\ 1\ 0\ 0}\rangle \oplus \mathbf{p_1},}{\Gamma \rhd \texttt{call } f : A} \quad (\text{VCALL})$$

Specific Features of the Program Logic

- Unusual rules for **mutually recursive methods** and for **parameter adaptation** in method invocations

$$\frac{(\Gamma, \mathsf{e} : A) \;\; goodContext}{\rhd\, \mathsf{e} : A} \qquad (\textsc{Mutrec})$$

$$\frac{(\Gamma, c \diamond m(\overline{a}) : MS \; c \;\; m \;\; \overline{a}) \;\; goodContext}{\rhd\, c \diamond m(\overline{b}) : MS \; c \;\; m \;\; \overline{b}} \qquad (\textsc{Adapt})$$

- Proof via admissible Cut rule, **no extra derivation system**
- Global specification table *MS*, *goodContext* relates entries in *MS* to the method bodies

Example: Insertion sort

Specification:

$$
\begin{array}{lcl}
insSpec & \equiv & MS \text{ List ins } [a_1, a_2] = \\
& & \quad \lambda\ E\ h\ h'\ v\ p\ .\forall\ i\ r\ n\ X\ . \\
& & \quad\ (E\langle a_1\rangle = i \wedge E\langle a_2\rangle = \text{Ref } r \wedge h, r \models_X n \\
& & \quad\ \longrightarrow |dom\,(h)| + 1 = |dom\,(h')| \wedge \\
& & \quad\quad p \leq \ldots) \\
sortSpec & \equiv & MS \text{ List sort } [a] = \\
& & \quad \lambda\ E\ h\ h'\ v\ p\ .\forall\ i\ r\ n\ X\ . \\
& & \quad\ (\ E\langle a\rangle = \text{Ref } r \wedge h, r \models_X n\ \longrightarrow |dom\,(h)| = |dom\,(h')|\ \wedge\ p \leq \ldots)
\end{array}
$$

Lemma:

$$insSpec\ \wedge\ sortSpec \longrightarrow\ \rhd \text{List} \diamond sort([xs]) : MS \text{ List } sort\ [xs]$$

Discussion of the Program Logic

- Expressive logic for correctness and resource consumption
- Logic proven **sound and complete**
- Termination built on top of a logic for partial correctness
- Less suited for immediate program verification: not fully automatic (case-splits, $\exists$-instantiations,...), verification conditions large and complex
- Continue abstraction: loop unfolding in op. semantics $\longrightarrow$ invariants in general program logics $\longrightarrow$ specific logic for interesting (resource-)properties
- Aim: exploit structure of Camelot compilation (freelist) and program analysis

$$\begin{array}{rcl} \text{List.ins} & : & 1, \mathbf{IL}(0) \rightarrow \mathbf{L}(0), 0 \\ \text{List.sort} & : & 0, \mathbf{L}(0) \rightarrow \mathbf{L}(0), 0 \end{array}$$

Heap Space Logic (LFD-assertions)

- Translation of Hofmann-Jost type system to Grail, types interpreted as relating initial to final freelist
- Fixed assertion format $[\![U, n, [\Delta] \blacktriangleright T, m]\!]$

$$\begin{array}{rcl} \text{List.ins} & : & [\![\{a, l\}, 1, [a \mapsto \mathbf{I}, l \mapsto \mathbf{L}(0)] \blacktriangleright \mathbf{L}(0), 0]\!] \\ \text{List.sort} & : & [\![\{l\}, 0, [l \mapsto \mathbf{L}(0)] \blacktriangleright \mathbf{L}(0), 0]\!] \end{array}$$

- LFD types express space requirements for datatype constructors, numbers n, m refer to the freelist length
- Semantic definition by expansion into core bytecode logic, derived proof rules using linear affine context management
- Dramatic reduction of VC complexity!

Semantic interpretation of $[\![U, n, [\Delta] \blacktriangleright T, m]\!]$

$$
\begin{array}{l}
[\![U, n, [\Delta] \blacktriangleright T, m]\!] \equiv \\
\;\lambda\ E\ h\ h'\ v\ p. \\
\quad \forall\ F\ N.\quad (regionsExist(U, \Delta, h, E) \wedge regionsDistinct(U, \Delta, h, E) \wedge \\
\qquad\qquad\quad freelist(h, F, N) \wedge distinctFrom(U, \Delta, h, E, F)) \\
\qquad\qquad\quad \longrightarrow \\
\qquad\qquad\quad (\exists\ R\ S\ M\ G.\ \ v, h' \models_T R, S\ \wedge freelist(h', G, M) \wedge R \cap G = \emptyset\ \wedge \\
\qquad\qquad\qquad\qquad Bounded((R \cup G), F, U, \Delta, h, E) \wedge modified(F, U, \Delta, h, E, \\
\qquad\qquad\qquad\qquad sizeRestricted(n, N, m, S, M, U, \Delta, h, E) \wedge dom\, h = dom\, h')
\end{array}
$$

- Formulae defined by BC expansion:

$$
\begin{array}{l}
regionsDistinct(U, \Delta, h, E) \equiv \\
\quad \forall\ x\ y\ R_x\ R_y\ S_x\ S_y. \\
\qquad (\{x, y\} \subseteq U \cap dom\,\Delta \wedge x \neq y \wedge E\langle x\rangle, h \models_{\Delta(x)} R_x, S_x \wedge E\langle y\rangle, h \models_{\Delta(y)} R_y, S_y) \\
\qquad \longrightarrow R_x \cap R_y = \emptyset \\
sizeRestricted(n, N, m, S, M, U, \Delta, h, E) \equiv \\
\quad \forall\ q\ C.\ Size(E, h, U, \Delta, C) \wedge n + C + q \leq N \longrightarrow m + S + q \leq M
\end{array}
$$

- **You don't want to read this — and you don't need to!**

Proof System

Proof system with linear inequalities and linear affine type system (U, Δ) that guarantees benign sharing;

$$\frac{\Delta(x) = T \quad n \leq m}{\Gamma \rhd \texttt{var } x : [\![\{x\}, m, [\Delta] \blacktriangleright T, n]\!]} \qquad (\textsc{Var})$$

$$\frac{\begin{array}{ll} \Gamma \rhd e_1 : [\![U_1, n, [\Delta] \blacktriangleright T_1, m]\!] & \Gamma \rhd e_2 : [\![U_2, m, [\Delta, x \mapsto T_1] \blacktriangleright T_2, k]\!] \\ U_1 \cap (U_2 \setminus \{x\}) = \emptyset & T_1 = \mathbf{L}(\_) \end{array}}{\Gamma \rhd \texttt{let } x = e_1 \texttt{ in } e_2 : [\![U_1 \cup (U_2 \setminus \{x\}), n, [\Delta] \blacktriangleright T_2, k]\!]} \qquad (\textsc{Let})$$

$$\frac{\Delta(x) = \mathbf{L}(k) \quad l = n + k \quad \Gamma \rhd e : [\![U, l, [\Delta, t \mapsto \mathbf{L}(k)] \blacktriangleright T, m]\!] \quad x \notin U \setminus \{t\}}{\Gamma \rhd \texttt{let } t = x.TL \texttt{ in } e : [\![(U \setminus \{t\}) \cup \{x\}, n, [\Delta] \blacktriangleright T, m]\!]} \qquad (\textsc{LetTL})$$

Note: Linearity relaxed in rules for compiled `match`-expressions

Discussion of the Heap Space Logic

- ☺ Exploit program structure and compiler analysis: most effort done once (in soundness proofs), application straight-forward
- ☺ "Classic PCC": independence of derived logic from Isabelle (no higher-order predicates, certifying constraint logic programming)
- ☺ "Foundational PCC": can unfold back to core logic and operational semantics if desired
- ☹ Generalisation to all Camelot datatypes needed
- ☹ Soundness proofs non-trivial, and challenging to generalise to more liberal sharing disciplines

Certificate Generation

Goal: Automatically generate proofs from high-level types and inferred heap consumption.

Approach: Use inferred space bounds as invariants. Use powerful Isabelle tactics to automatically prove a statement on heap consumption in the heap logic.

Example certificate (for list append):

$\Gamma \rhd$ snd (methtable Append append) : SPEC append
by (Wp append_pdefs)

$\rhd$Append.append([RNarg x0, RNarg x1]) : sMST Append append [RNarg x0, RNarg
by (fastsimp intro: Context_good GCInvs simp: ctxt_def)

Summary

MRG works towards **resource-safe global computing**:

- **check resource consumption** before executing downloaded code;
- **automatically generate certificate** out of a Camelot type.

Components of the picture

- Proof-Carrying-Code infrastructure
- Inference for space consumption in Camelot
- Specialised derived assertions on top of a general program logic for Grail
- Certificate = proof of a derived assertion
- Certificate generation from high-level types

Further Reading

- David Aspinall, Stephen Gilmore, Martin Hofmann, Donald Sannella and Ian Stark, *Mobile Resource Guarantees for Smart Devices* in CASSIS04 — Construction and Analysis of Safe, Secure, and Interoperable Smart Devices, LNCS 3362, 2005. `http://groups.inf.ed.ac.uk/mrg/publications/mrg/cassis2004.p`
- David Aspinall and Lennart Beringer and Martin Hofmann and Hans-Wolfgang Loidl and Alberto Momigliano, *A Program Logic for Resource Verification*, in TPHOLs2004 — International Conference on Theorem Proving in Higher Order Logics, Utah, LNCS 3223, 2004.
- Martin Hofmann, Steffen Jost, *Static Prediction of Heap Space Usage for First-Order Functional Programs*, in POPL'03 — Symposium on Principles of Programming Languages, New Orleans, LA, USA, Jan 2003.

Further Reading

- K. Crary and S. Weirich, *Resource Bound Certification* in POPL'00 — Symposium on Principles of Programming Languages, Boston, USA, 2000.
`http://www-2.cs.cmu.edu/ crary/papers/1999/res/res.ps.gz`
- Gilles Barthe, Mariela Pavlova, Gerardo Schneider, *Precise analysis of memory consumption using program logics* in International Conference on Software Engineering and Formal Methods (SEFM 2005), 7–9 September 2005, Koblenz, Germany.
`http://www-sop.inria.fr/everest/soft/Jack/doc/papers/gmg05.p`

Approaches to Certificate Generation

One of the main problems of PCC is how to generate the proofs.

Different approaches are:

- Type system (Necula, MRG)
- Abstract interpretation (Certified A.I.)
- Model checking

Abstract Interpretation Based

Certified abstract interpretation is a technique for extracting a static analyser from the **constructive proof of its semantic correctness**, producing at the same time an analyser and a proof object certifying its semantic correctness.

Abstract Interpretation Based

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Main advantages

- ☺ additional flexibility
- ☺ foundational nature

Proof Checkers

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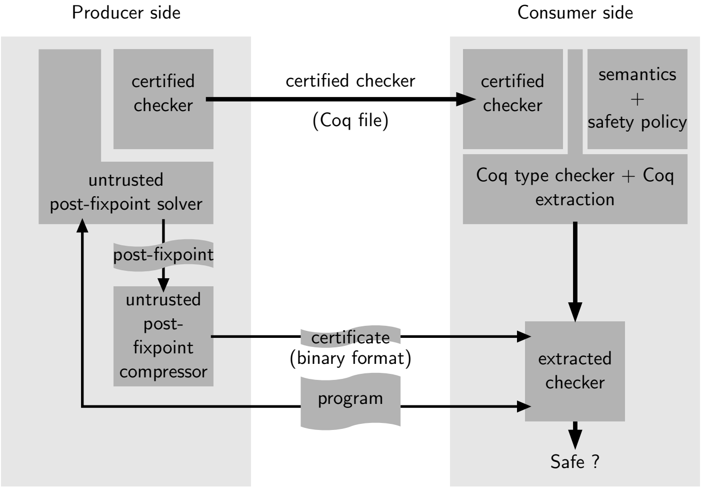
Special purpose proof checkers (e.g. for Java bytecode verification) are fast and small, with small certificates, but in general not as trustworthy.

Idea: By using a PCC approach on the proof checker itself, we can maintain a trustworthy core system and simplify certificate generation.
$\Longrightarrow$ **proof carrying proof checker**.

PCC with Certified Abstract Interpretation



Trusted Code Base

The trusted code base comprises:

- a formalisation of the semantics of the language;
- a (semantic) formalisation of the security policy;
- a core proof checker to be applied on a proof carrying proof checker;

The abstract interpretation machinery annotates a program with properties at program points and finds a fixed point.

Components of this PCC architecture

Uses generic abstract interpretation machinery:

- Abstract value: specific to the analysis
- Complete certificate = set of program annotations, encoding information on the abstract values
- Abstract state = mapping of program points to abstract memories
- Validation = check that all annotations are fulfilled and that the annotations imply the security policy
- The proof checker performs fixpoint iteration over the abstract domain, until the annotations are met

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All this is implemented in the Coq theorem prover.

Reducing Certificate Size

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Reduce certificate size: Program only sparsely annotated; a reconstructions algorithm is run at consumer side to obtain all annotations.

Reduce validation time: Attach to the certificate a strategy that guides the reconstruction algorithm (e.g. where is it safe to drop annotations during reconstruction). Similar to “oracle strings”.

Efficiency of Validation

Program	checking time (sec)	analyser/checker (no. of constaints)
BubbleSort	0.015	440/110
HeapSort	0.050	8001/381
QuickSort	0.060	8910/405
Convolution	0.010	460/92
FloydWarshall	0.020	23114/163
PolynomProduct	0.010	150669/133

Size of Certificates

Program	.java	.class	complete fixpoint	compr'd fixpoint	bin cert	bin cert/ .class
BubbleSort	440	528	3640	182	44	8.3%
HeapSort	1044	858	17352	332	63	7.3%
QuickSort	1078	965	25288	629	158	16.4%
Convolution	378	542	2942	195	52	9.6%
FloydWarshall	417	596	7180	346	134	22.5%
PolynomProduct	509	604	5366	308	87	14.4%

Summary

PCC is a powerful, general mechanism for providing safety guarantees for mobile code.

It provides these guarantees without resorting to a trust relationship.

It uses techniques from the areas of type-systems, program verification and logics.

It is a very active research area at the moment.

Current Trends

Using formal methods to check specific program properties.

- Program logics as the basic language for doing these checks attract renewed interest in PCC.
- A lot of work on program logics for low-level languages.
- Immediate applications for smart cards and embedded systems.

Future Directions

Embedded Systems as a domain for formal methods.

- Some of these systems have strong security requirements.
- Formal methods are used to check these requirements.
- Model checking is a very active area for automatically checking properties.

Links to other areas

Checking program properties is closely related to inferring quantitative information.

- **Static analyses** deal with extracting quantitative information (e.g. resource consumption)
- A lot of research has gone into making these techniques efficient.
- **Model checking** can deal with a larger class of problems (e.g. specifying safety conditions in a system)
- Just recently these have become efficient enough to be used for main stream programming.

Reading List:

`http://www.tcs.ifi.lmu.de/~hwloidl/PCC/reading.html`