Introduction

MPRI 2–6: Abstract Interpretation,
application to verification and static analysis

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year 2015–2016

course 01
16 September 2015

Motivating program verification

The cost of software failure

- Patriot MIM-104 failure, 25 February 1991
  (death of 28 soldiers[1])
- Ariane 5 failure, 4 June 1996
  (cost estimated at more than 370 000 000 US$[2])
- Toyota electronic throttle control system failure, 2005
  (at least 89 death[3])
- Heartbleed bug in OpenSSL, April 2014
- Stagefright bug in Android, Summer 2015
  (multiple array overflows in 900 million devices, some exploitable)
- economic cost of software bugs is tremendous[4]

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[1] R. Skeel. "Roundoff Error and the Patriot Missile". SIAM News, volume 25, nr 4.

[2] M. Dowson. "The Ariane 5 Software Failure". Software Engineering Notes 22 (2): 84, March 1997.

[3] CBSNews. Toyota "Unintended Acceleration" Has Killed 89. 20 March 2014.

[4] NIST. Software errors cost U.S. economy $59.5 billion annually. Tech. report, NIST Planning Report, 2002.

Zoom on: Ariane 5, Flight 501

Maiden flight of the Ariane 5 Launcher, 4 June 1996.

Zoom on: Ariane 5, Flight 501

40s after launch. . .

Zoom on: Ariane 5, Flight 501

Cause: software error[5]

- arithmetic overflow in unprotected data conversion from 64-bit float to 16-bit integer types[6]

```
P_M_DERIVE(T_ALG.E_BH) :=
  UC_16S_EN_16NS (TDB.T_ENTIER_16S
    ((1.0/C_M_LSB_BH) * G_M_INFO_DERIVE(T_ALG.E_BH)));
```

- software exception not caught
  $\Longrightarrow$ computer switched off
- all backup computers run the same software
  $\Longrightarrow$ all computers switched off, no guidance
  $\Longrightarrow$ rocket self-destructs

5 J.-L. Lions et al., Ariane 501 Inquiry Board report.

6 J.-J. Levy. Un petit bogue, un grand boum. Séminaire du Département d'informatique de l'ENS, 2010.

How can we avoid such failures?

- Choose a safe programming language.

  C (low level) / Ada, Java (high level)

- Carefully design the software.

  many software development methods exist

- Test the software extensively.

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  yet, Ariane 5 software is written in Ada

- Carefully design the software.

  many software development methods exist

  yet, critical embedded software follow strict development processes

- Test the software extensively.

  yet, the erroneous code was well tested. . . on Ariane 4

$\Longrightarrow$ **not sufficient!**

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  C (low level) / Ada, Java (high level)

  yet, Ariane 5 software is written in Ada

- Carefully design the software.

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  yet, critical embedded software follow strict development processes

- Test the software extensively.

  yet, the erroneous code was well tested. . . on Ariane 4

$\Longrightarrow$ **not sufficient!**

We should use **formal methods.**

provide rigorous, mathematical insurance

Proving program properties

Invariants and programs

```
assume X in [0,1000];

I := 0;

while I < X do

    I := I + 2;


assert I in [0,?]
```

Goal: find a bound property, sufficient to express the absence of overflow

[7] R. W. Floyd. "Assigning meanings to programs". In Proc. Amer. Math. Soc. Symposia in Applied Mathematics, vol. 19, pp. 19–31, 1967.

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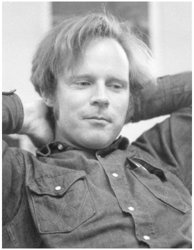
Invariants and programs

```
assume X in [0,1000];
```
$\{X \in [0,1000]\}$
```
I := 0;
```
$\{X \in [0,1000], I = 0\}$
```
while I < X do
```
$\{X \in [0,1000], I \in [0,998]\}$
```
    I := I + 2;
```
$\{X \in [0,1000], I \in [2,1000]\}$

$\{X \in [0,1000], I \in [0,1000]\}$
```
assert I in [0,1000]
```



Robert Floyd[7]

invariant: property true of all the executions of the program

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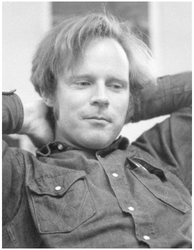
Invariants and programs

```
assume X in [0,1000];
```
$\{X \in [0, 1000]\}$
```
I := 0;
```
$\{X \in [0, 1000], I = 0\}$
```
while I < X do
```
  $\{X \in [0, 1000], I \in [0, 998]\}$
```
    I := I + 2;
```
  $\{X \in [0, 1000], I \in [2, 1000]\}$

$\{X \in [0, 1000], I \in [0, 1000]\}$
```
assert I in [0,1000]
```



Robert Floyd[7]

invariant: property true of all the executions of the program

issue: if $I = 997$ at a loop iteration, $I = 999$ at the next

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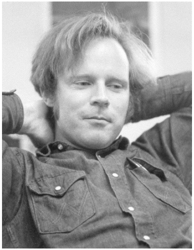
```
assume X in [0,1000];
```
$\{X \in [0, 1000]\}$
```
I := 0;
```
$\{X \in [0, 1000], I = 0\}$
```
while I < X do
```
$\{X \in [0, 1000], I \in \{0, 2, \ldots, 996, 998\}\}$
```
    I := I + 2;
```
$\{X \in [0, 1000], I \in \{2, 4, \ldots, 998, 1000\}\}$

$\{X \in [0, 1000], I \in \{0, 2, \ldots, 998, 1000\}\}$
```
assert I in [0,1000]
```



Robert Floyd[7]

inductive invariant: invariant that can be proved to hold by induction on loop iterates

(if $I \in S$ at a loop iteration, then $I \in S$ at the next loop iteration)

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Logics and programs

$$\frac{}{\{P[e/X]\}\ \texttt{X := e}\ \{P\}} \qquad \frac{\{P\}\ \texttt{C}_1\ \{R\} \quad \{R\}\ \texttt{C}_2\ \{Q\}}{\{P\}\ \texttt{C}_1;\texttt{C}_2\ \{Q\}}$$

$$\frac{\{P \,\&\, b\}\ \texttt{C}\ \{P\}}{\{P\}\ \texttt{while b do C}\ \{P \,\&\, \neg b\}}$$

...

Tony Hoare[8]

- sound logic to prove program properties, (rel.) complete
- proofs can be partially automated (at least proof checking)
  (e.g., using proof assistants: Coq, PVS, Isabelle, HOL, etc.)

[8] C. A. R. Hoare. "An Axiomatic Basis for Computer Programming". Commun. ACM 12(10): 576–580 (1969).

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$$\ldots$$

Tony Hoare[8]

- sound logic to prove program properties, (rel.) complete
- proofs can be partially automated (at least proof checking)
  (e.g., using proof assistants: Coq, PVS, Isabelle, HOL, etc.)
- requires annotations and interaction with a prover
  even manual annotation is not practical for large programs

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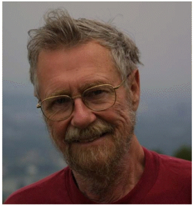
A calculs of program properties

$$wlp(\texttt{X} := \texttt{e}, P) \stackrel{\text{def}}{=} P[e/X]$$

$$wlp(\texttt{C}_1; \texttt{C}_2, P) \stackrel{\text{def}}{=} wlp(\texttt{C}_1, wlp(\texttt{C}_2, P))$$

$$wlp(\texttt{while e do C}, P) \stackrel{\text{def}}{=} I \wedge ((e \wedge I) \implies wlp(\texttt{C}, I)) \wedge ((\neg e \wedge I) \implies P)$$



Edsger W. Dijkstra[9]

- predicate transformer semantics
  propagate predicates on states through the program
- weakest (liberal) precondition
  backwards, from property to prove to condition for program correctness
- calculs that can be mostly automated

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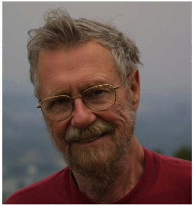
$$I \wedge ((e \wedge I) \implies wlp(\texttt{C}, I)) \wedge ((\neg e \wedge I) \implies P)$$



Edsger W. Dijkstra[9]

- predicate transformer semantics
  propagate predicates on states through the program
- weakest (liberal) precondition
  backwards, from property to prove to condition for program correctness
- calculs that can be mostly automated, except for:
  - user annotations for inductive loop invariants
  - function annotations (modular inference)
- academic success: complex (functional) but local properties
- industry success: simple and local properties

[9] E. W. Dijkstra. "Guarded commands, nondeterminacy and formal derivation of programs". EWD472. Commun. ACM 18(8): 453-457 (1975).

Limit to automation

Computers, programs, data

$$O(P, D) \in \{yes, no, \bot\}$$

O

P

D

The computer O runs the program P on the data D and answers (yes,no)... or does not answer ($\bot$).

Computers, programs, data

$$O(P, D) \in \{yes, no, \bot\}$$

O

P

P'

Note that programs are also a kind of data!

They can be fed to other programs. (e.g., to compilers)

Static analysis

Static analyzer A.
Given a program P:

- $O(A, P) = yes \iff \forall D,\ O(P, D)$ is safe
- $O(A, P) \neq \perp$ (the static analysis always terminates)

Static analysis

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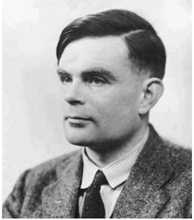
Given a program P:

- $O(A, P) = yes \iff \forall D,\ O(P, D)$ is safe
- $O(A, P) \neq \bot$ (the static analysis always terminates)

$\implies P$ is proved safe even before it is run!

Fundamental undecidability

There cannot exist a static analyzer A proving the termination of every terminating program P.



Alan Turing[10]

[10] A. M. Turing. "Computability and definability". The Journal of Symbolic Logic, vol. 2, pp. 153–163, (1937).

[11] H. G. Rice. "Classes of Recursively Enumerable Sets and Their Decision Problems." Trans. Amer. Math. Soc. 74, 358-366, 1953.

Fundamental undecidability

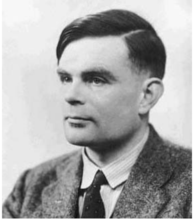
There cannot exist a static analyzer A proving the termination of every terminating program P.

Proof sketch:

$$A(P \cdot D) : O(A, P \cdot D) = \left| \begin{array}{l} \textit{yes} \text{ if } O(P,D) \neq \bot \\ \textit{no} \text{ otherwise} \end{array} \right.$$

$A'(X)$: `while A(X·X) do` *nothing* ; *no*

do we have $O(A', A') = \bot$ or $\neq \bot$? neither!
$\Longrightarrow$ A cannot exist



Alan Turing[10]

All "interesting" properties are undecidable![11] 

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Approximation

Approximate static analysis

An approximate static analyzer A always answers in finite time ($\neq \bot$):

- either *yes*: the program P is definitely safe (soundness)
- either *no*: I don't know (incompleteness)

Sufficient to prove the safety of (some) programs.
Fails on infinitely many programs...

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An approximate static analyzer A always answers in finite time ($\neq \bot$):

- either *yes*: the program P is definitely safe (soundness)
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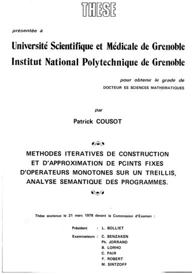
Sufficient to prove the safety of (some) programs.
Fails on infinitely many programs...

$\Longrightarrow$ We should adapt the analyzer A to

- a class of programs to verify, and
- a class of safety properties to check.

Abstract interpretation

Patrick Cousot[12]



THESE

présentée à

Université Scientifique et Médicale de Grenoble
Institut National Polytechnique de Grenoble

pour obtenir le grade de
DOCTEUR ES SCIENCES MATHEMATIQUES

par

Patrick COUSOT

METHODES ITERATIVES DE CONSTRUCTION
ET D'APPROXIMATION DE PCINTS FIXES
D'OPERATEURS MONOTONES SUR UN TREILLIS,
ANALYSE SEMANTIQUE DES PROGRAMMES.

Thèse soutenue le 21 mars 1978 devant la Commission d'Examen :

Président : L. BOLLIET

Examinateurs : C. BENZAKEN
Ph. JORRAND
B. LORHO
C. PAIR
F. ROBERT
M. SINTZOFF

General theory of the approximation and comparison of program semantics:

- unifies many existing semantics
- allows the definition of new static analyses that are correct by construction

[12] P. Cousot. "Méthodes itératives de construction et d'approximation de points fixes d'opérateurs monotones sur un treillis, analyse sémantique des programmes." Thèse És Sciences Mathématiques, 1978.

Abstract interpretation

$(\mathcal{S}_0)$

```
assume X in [0,1000];
```

$(\mathcal{S}_1)$

```
I := 0;
```

$(\mathcal{S}_2)$

```
while (S3) I < X do
    (S4)
    I := I + 2;
    (S5)
```

$(\mathcal{S}_6)$

program

Abstract interpretation

```
(S0)
assume X in [0,1000];
(S1)
I := 0;
(S2)
while (S3) I < X do
    (S4)
    I := I + 2;
    (S5)
(S6)
```

program

$\mathcal{S}_i \in \mathcal{D} = \mathcal{P}(\{\mathtt{I},\mathtt{X}\} \to \mathbb{Z})$

$$
\begin{array}{lll}
\mathcal{S}_0 = \{\, (i,x) \mid i,x \in \mathbb{Z} \,\} & = \top \\
\mathcal{S}_1 = \{\, (i,x) \in \mathcal{S}_0 \mid x \in [0,1000] \,\} & = F_1(\mathcal{S}_0) \\
\mathcal{S}_2 = \{\, (0,x) \mid \exists i,\ (i,x) \in \mathcal{S}_1 \,\} & = F_2(\mathcal{S}_1) \\
\mathcal{S}_3 = \mathcal{S}_2 \cup \mathcal{S}_5 & \\
\mathcal{S}_4 = \{\, (i,x) \in \mathcal{S}_3 \mid i < x \,\} & = F_4(\mathcal{S}_3) \\
\mathcal{S}_5 = \{\, (i+2,x) \mid (i,x) \in \mathcal{S}_4 \,\} & = F_5(\mathcal{S}_4) \\
\mathcal{S}_6 = \{\, (i,x) \in \mathcal{S}_3 \mid i \geq x \,\} & = F_6(\mathcal{S}_3)
\end{array}
$$

semantics

Concrete semantics $\mathcal{S}_i \in \mathcal{D} = \mathcal{P}(\{\mathtt{I},\mathtt{X}\} \to \mathbb{Z})$:

- strongest invariant (and an inductive invariant)
- not computable in general
- smallest solution of a system of equations

Abstract interpretation

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```
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```
while
```
$(\mathcal{S}_3)$
```
I < X do
```
$(\mathcal{S}_4)$
```
    I := I + 2;
```
$(\mathcal{S}_5)$

$(\mathcal{S}_6)$

program

$$\mathcal{S}_i^\sharp \in \mathcal{D}^\sharp$$

$$\begin{aligned}
\mathcal{S}_0^\sharp &= \top^\sharp \\
\mathcal{S}_1^\sharp &= F_1^\sharp(\mathcal{S}_0^\sharp) \\
\mathcal{S}_2^\sharp &= F_2^\sharp(\mathcal{S}_1^\sharp) \\
\mathcal{S}_3^\sharp &= S_2^\sharp \cup^\sharp \mathcal{S}_5^\sharp \\
\mathcal{S}_4^\sharp &= F_4^\sharp(\mathcal{S}_3^\sharp) \\
\mathcal{S}_5^\sharp &= F_5^\sharp(\mathcal{S}_4^\sharp) \\
\mathcal{S}_6^\sharp &= F_6^\sharp(\mathcal{S}_3^\sharp)
\end{aligned}$$

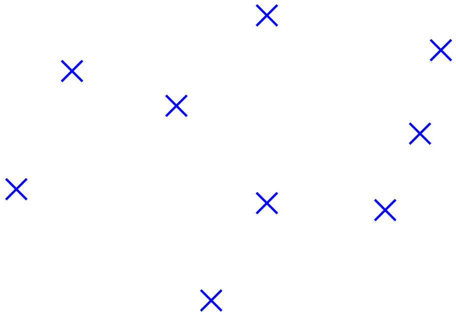
semantics

Abstract semantics $\mathcal{S}_i^\sharp \in \mathcal{D}^\sharp$:

- $\mathcal{D}^\sharp$ is a subset of properties of interest (approximation) with a machine representation
- $F^\sharp : \mathcal{D}^\sharp \to \mathcal{D}^\sharp$ over-approximates the effect of $F : \mathcal{D} \to \mathcal{D}$ in $\mathcal{D}^\sharp$

  (with effective algorithms)

Numeric abstract domain examples



concrete sets $\mathcal{D}$: $\{(0,3),(5.5,0),(12,7),\ldots\}$

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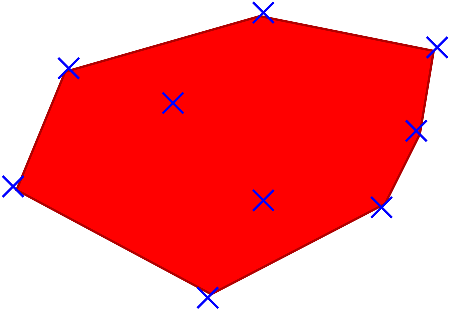

concrete sets $\mathcal{D}$: $\{(0,3),(5.5,0),(12,7),\ldots\}$

abstract polyhedra $\mathcal{D}^\sharp_p$: $6X + 11Y \geq 33 \wedge \cdots$

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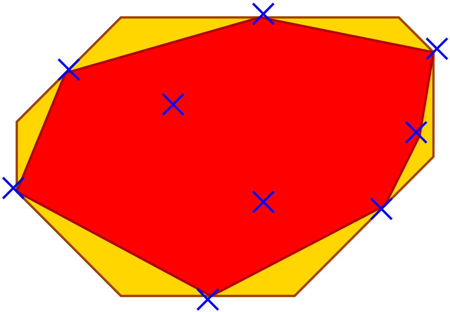
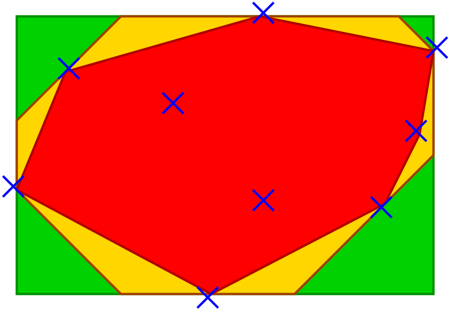


concrete sets $\mathcal{D}$: $\{(0,3),(5.5,0),(12,7),\ldots\}$

abstract polyhedra $\mathcal{D}_p^\sharp$: $6X + 11Y \geq 33 \wedge \cdots$

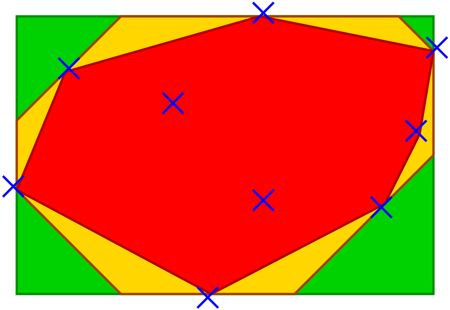
abstract octagons $\mathcal{D}_o^\sharp$: $X + Y \geq 3 \wedge Y \geq 0 \wedge \cdots$

Numeric abstract domain examples



concrete sets $\mathcal{D}$:	$\{(0,3),(5.5,0),(12,7),\ldots\}$
abstract polyhedra $\mathcal{D}^\sharp_p$:	$6X + 11Y \geq 33 \wedge \cdots$
abstract octagons $\mathcal{D}^\sharp_o$:	$X + Y \geq 3 \wedge Y \geq 0 \wedge \cdots$
abstract intervals $\mathcal{D}^\sharp_i$:	$X \in [0,12] \wedge Y \in [0,8]$

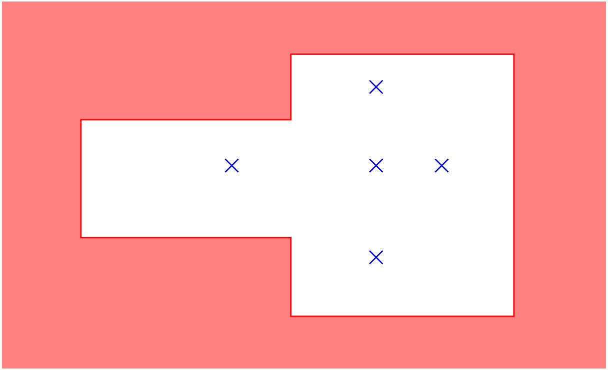
Numeric abstract domain examples



concrete sets $\mathcal{D}$:	$\{(0,3),(5.5,0),(12,7),\dots\}$	not computable
abstract polyhedra $\mathcal{D}^\sharp_p$:	$6X + 11Y \geq 33 \wedge \cdots$	exponential cost
abstract octagons $\mathcal{D}^\sharp_o$:	$X + Y \geq 3 \wedge Y \geq 0 \wedge \cdots$	cubic cost
abstract intervals $\mathcal{D}^\sharp_i$:	$X \in [0,12] \wedge Y \in [0,8]$	linear cost

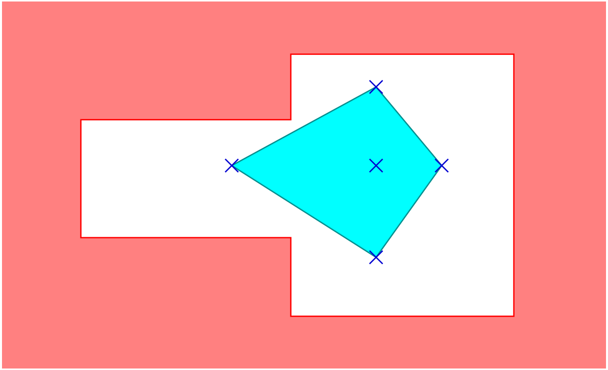
Trade-off between cost and expressiveness / precision

Correctness proof and false alarms



The program is correct (blue $\cap$ red $= \emptyset$).

Correctness proof and false alarms



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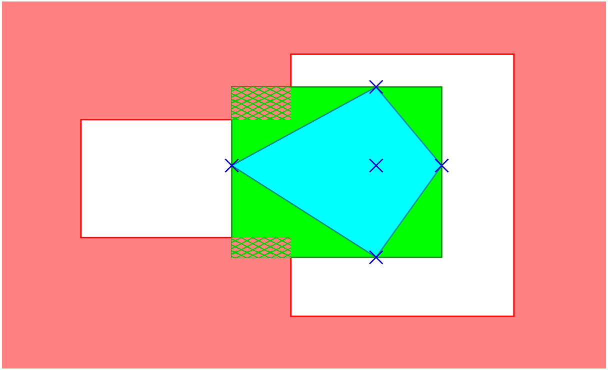
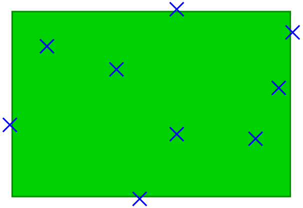
The polyhedra domain can prove the correctness (cyan $\cap$ red $= \emptyset$).

Correctness proof and false alarms



The program is correct (blue $\cap$ red $= \emptyset$).
The polyhedra domain can prove the correctness (cyan $\cap$ red $= \emptyset$).
The interval domain cannot (green $\cap$ red $\neq \emptyset$, false alarm).

Numeric abstract domain examples (cont.)



abstract semantics $F^\sharp$ in the interval domain $\mathcal{D}_i^\sharp$

- $I \in \mathcal{D}_i^\sharp$ is a pair of bounds $(\ell, h) \in \mathbb{Z}^2$ (for each variable) representing an interval $[\ell, h] \subseteq \mathbb{Z}$
- `I:=I+2`: $(\ell, h) \mapsto (\ell+2, h+2)$
- $\cup^\sharp$: $(\ell_1, h_1) \cup^\sharp (\ell_2, h_2) = (\min(\ell_1, \ell_2), \max(h_1, h_2))$
- ...

Resolution by iteration and extrapolation

Challenge: the equation system is recursive: $\vec{\mathcal{S}}^\sharp = \vec{F}^\sharp(\vec{\mathcal{S}}^\sharp)$.

Solution: resolution by iteration: $\vec{\mathcal{S}}^{\sharp\,0} = \emptyset^\sharp$, $\vec{\mathcal{S}}^{\sharp\,i+1} = \vec{F}^\sharp(\vec{\mathcal{S}}^{\sharp\,i})$.

e.g., $\mathcal{S}_3^\sharp$: $\texttt{I} \in \emptyset$, $\texttt{I} = 0$, $\texttt{I} \in [0,2]$, $\texttt{I} \in [0,4]$, $\ldots$, $\texttt{I} \in [0,1000]$

Resolution by iteration and extrapolation

Challenge: the equation system is recursive: $\vec{\mathcal{S}}^\sharp = \vec{F}^\sharp(\vec{\mathcal{S}}^\sharp)$.

Solution: resolution by iteration: $\vec{\mathcal{S}}^{\sharp\,0} = \emptyset^\sharp$, $\vec{\mathcal{S}}^{\sharp\,i+1} = \vec{F}^\sharp(\vec{\mathcal{S}}^{\sharp\,i})$.

e.g., $\mathcal{S}_3^\sharp$: $\texttt{I} \in \emptyset$, $\texttt{I} = 0$, $\texttt{I} \in [0, 2]$, $\texttt{I} \in [0, 4]$, ..., $\texttt{I} \in [0, 1000]$

Challenge: infinite or very long sequence of iterates in $\mathcal{D}^\sharp$

Solution: extrapolation operator $\triangledown$

e.g., $[0, 2] \triangledown [0, 4] = [0, +\infty[$

- remove unstable bounds and constraints
- ensures the convergence in finite time
- inductive reasoning (through generalisation)

Resolution by iteration and extrapolation

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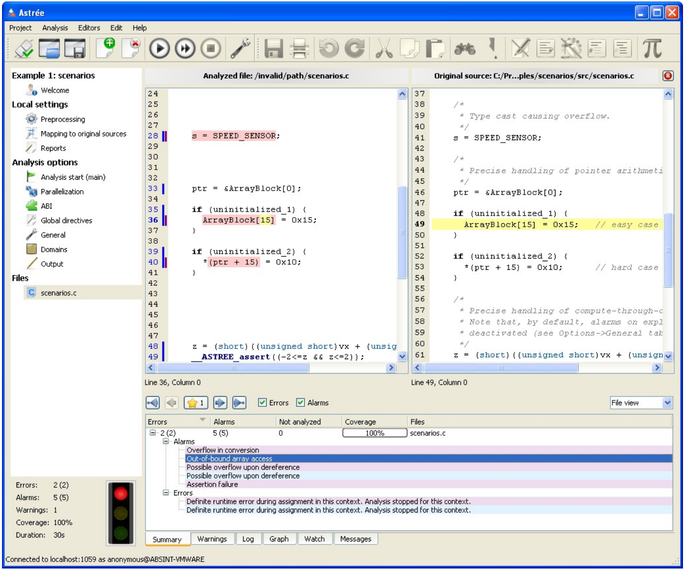
$\Longrightarrow$ effective solving method $\longrightarrow$ static analyzer!

Other uses of abstract interpretation

- Analysis of dynamic memory data-structures (*shape analysis*).
- Analysis of parallel, distributed, and multi-thread programs.
- Analysis of probabilistic programs.
- Analysis of biological systems.
- Security analysis (*information flow*).
- Termination analysis.
- Cost analysis.
- Analyses to enable compiler optimisations.
- ...

Some static analysis tools based on Abstract Interpretation

The Astrée static analyzer



The Astrée static analyzer

Analyseur statique de programmes temps-réels embarqués

(static analyzer for real-time embedded software)

- developed at ENS

  B. Blanchet, P. Cousot, R. Cousot, J. Feret,
  L. Mauborgne, D. Monniaux, A. Miné, X. Rival

- industrialized and made commercially available by AbsInt

Astrée

`www.astree.ens.fr`

AbsInt

`www.absint.com`

The Astrée static analyzer

Specialized:

- for the analysis of run-time errors

  (arithmetic overflows, array overflows, divisions by 0, etc.)

- on embedded critical C software

  (no dynamic memory allocation, no recursivity)

- in particular on control / command software

  (reactive programs, intensive floating-point computations)

- intended for validation

  (analysis does not miss any error and tries to minimise false alarms)

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Approximately 40 abstract domains are used at the same time:

- numeric domains (intervals, octagons, ellipsoids, etc.)
- boolean domains
- domains expressing properties on the history of computations

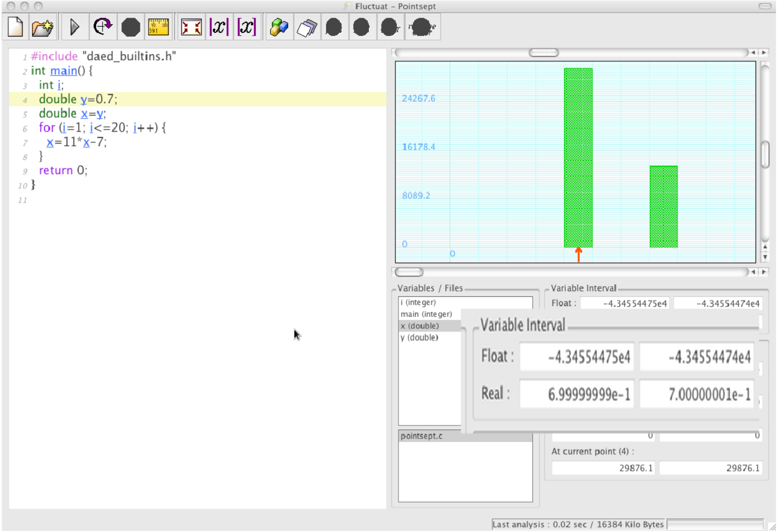
Astrée applications

Airbus A340-300 (2003)

Airbus A380 (2004)

- size: from 70 000 to 860 000 lines of C
- analysis time: from 45mn to $\simeq$40h
- 0 alarm: proof of absence of run-time error

Fluctuat


Fluctuat – Pointsept
#include "daed_builtins.h"
int main() {
int i;
double y=0.7;
double x=y;
for (i=1; i<=20; i++) {
x=11*x-7;
}
return 0;
}
24267.6
16178.4
8089.2
0
0
Variables / Files
i (integer)
main (integer)
x (double)
y (double)
pointsept.c
Variable Interval
Float : -4.34554475e4 -4.34554474e4
Real : 6.99999999e-1 7.00000001e-1
At current point (4) :
29876.1 29876.1
Last analysis : 0.02 sec / 16384 Kilo Bytes

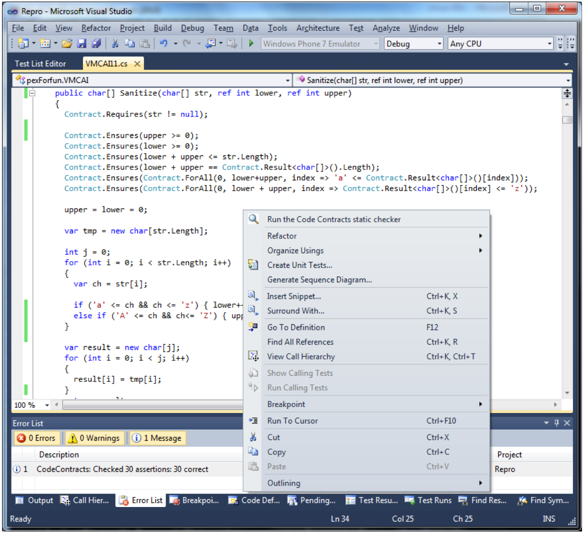
Fluctuat

Static analysis of the **accuracy of floating-point computations**:

- bound the range of variables
- bound the rounding errors wrt. real computation
- track the origin of rounding errors
  (which operation contributes to most error,
  target for improvements)
- uses specific abstract domains
  (affine arithmetic, zonotopes)

- developed at CEA-LIST (E. Goubault, S. Putot)
- industrial use (Airbus)

Clousot: CodeContract static checker



Clousot: CodeContract static checker

CodeContracts:

- assertion language for .NET (C#, VB, etc.)

  (pre-conditions, post-conditions, invariants)
- dynamic checking

  (insert run-time checks)
- static checking

  (modular abstract interpretation)
- automatic inference

  (abstract interpretation to infer necessary preconditions backwards)

- developed at Microsoft Research (M. Fahndrich, F. Logozzo)
- part of .NET Framework 4.0
- integrated to Visual Studio