- Types are a static system guaranteed by your program
- Types serve as evidence of a particular property, that relates to the structure of information
- For the lambda calculus, and base values, the only structure to be had is lambdas
- Type systems are designed to ensure certain static properties of the language. These properties can be relatively superficial, or fairly involved
- The simply-typed lambda calculus is one specific type system for the lambda calculus that models all of the things that could “go wrong” at the type level
- Start by type system for IfArith
Higher-order contract systems track program labels alongside contracts to *properly assign blame* when failure occurs.

“I take in a positive and produce a positive.”

(define/contract (fib x)
  (-> positive? positive?)
  (cond
    [(= x 0) 1]
    [(= x 1) 1]
    [else (+ (fib (- x 1)) (fib (- x 2)))]))

Welcome to DrRacket, version 7.2 [3m].
Language: racket, with debugging; memory limit: 128 MB.
> (fib 2)
2
>
When I mess up

(define/contract (fib x)
  (-> positive? positive?)
  (cond
    [(= x 0) 1]
    [(= x 1) 1]
    [else (+ (fib (- x 1)) (fib (- x 2)))])))

> (fib -2)
  fib: contract violation
  expected: positive?
  given:  -2
  in:  the 1st argument of
        (-> positive? positive?)
  contract from: (function fib)
  blaming from: (function module)
  (assuming the contract is correct)
  at:  unsaved-editor:3.18
When I mess up

(define/contract (fib x)
  (-> positive? positive?))
(cond
  [(= x 0) 1]
  [(= x 1) 1]
  [else (+ (fib (- x 1)) (fib (- x 2)))]))

> (fib -2)

fib: contract violation
  expected: positive?
  given: -2
  in: the 1st argument of
    (-> positive? positive?)
  contract from: (function fib)
  blaming: anonymous-module
    (assuming the contract is correct)
  at: unsaved-editor:3.18

Racket blames me
(anonymous-module)
When \texttt{fib} messes up

\begin{verbatim}
(define/contract (fib x)
  (-> positive? positive?)
  (cond
    [(= x 0) -200]
    [(= x 1) 1]
    [else (+ (fib (- x 1)) (fib (- x 2)))]))
\end{verbatim}

Welcome to \texttt{DrRacket}, version 7.2 [3m].
Language: \texttt{racket}, with debugging; memory limit: 128 MB.

> (fib 20)

\texttt{fib: broke its own contract}
\texttt{promised: positive?}
\texttt{produced: }-829435
\texttt{in: the range of}
  \texttt{(-> positive? positive?)}
\texttt{contract from: (function fib)}
\texttt{blaming: (function fib)}
  \texttt{(assuming the contract is correct)}
\texttt{at: unsaved-editor:3.18}

Racket blames \texttt{fib}
Earlier...

Note that contracts are checked at **runtime**

*(Not compile time!)*

But sometimes we want to know our program won’t break **before** it runs!
Type Systems

A type system assigns each source fragment with a given type: a specification of how it will behave.

Type systems include rules, or judgements that tells us how we compositionally build types for larger fragments from smaller fragments.

The goal of a type system is to rule out programs that would exhibit run time type errors!
A type system for STLC
(Simply-Typed Lambda Calculus)

e ::= (lambda (x) e)
    | (e e)
    | ((prim e) e)
    | x
    | n

prim ::= + | * | ...
**Term Syntax**

\[
e ::= (\text{lambda} \ (x) \ e) \\
\quad | \ (e \ e) \\
\quad | \ ((\text{prim} \ e) \ e) \\
\quad | \ x \\
\quad | \ n
\]

\[
\text{prim} ::= + \ | \ * \ | \ ...
\]

**Type Syntax**

\[
t ::= \text{nat} \\
\quad | \ \text{bool} \\
\quad | \ t \to \ t
\]
Term Syntax

e ::= (lambda (x) e)  
| (e e)  
| ((prim e) e)  
| x  
| n

prim ::= + | * | ...

Type Syntax

t ::= nat  
| bool  
| t -> t

Function Types
**Term Syntax**

\[
\begin{align*}
e & ::= (\text{lambda} \ (x) \ e) \\
    & | \ (e \ e) \\
    & | \ ((\text{prim} \ e) \ e) \\
    & | \ x \\
    & | \ n
\end{align*}
\]

\[
\begin{align*}
\text{prim} & ::= + \\
    & | \ * \\
    & | \ ...
\end{align*}
\]

**Type Syntax**

\[
\begin{align*}
t & ::= \text{nat} \\
    & | \ \text{bool} \\
    & | \ t \rightarrow \ t
\end{align*}
\]

Examples...

\[
\begin{align*}
(\text{int} \rightarrow \text{int}) \rightarrow \text{int} \\
\text{bool} \rightarrow \text{int} \\
\text{bool} \rightarrow (\text{int} \rightarrow \text{bool})
\end{align*}
\]
;; Expressions are ifarith, with several special builtins
(define (expr? e)
  (match e
    ;; Variables
    [([? symbol? x) #t]
    ;; Literals
    [(? bool-lit? b) #t]
    [(? int-lit? i) #t]
    ;; Applications
    [`(,(? expr? e0) ,(? expr? e1)) #t]
    ;; Annotated expressions
    [`(,(? expr? e) : ,(? type? t)) #t]
    ;; Annotated lambdas
    [`(lambda (,(? symbol? x) : ,(? type? t)) ,(? expr? e)) #t]])
A type system for STLC

Type rules are written in natural-deduction style
(Like our big-step operational semantics.)

Assumptions above the line  (No assumptions here.)

\[ n : \text{num} \]

Conclusions below the line

\[ \text{Const} \]

\[
\begin{align*}
e ::= & \text{(lambda } (x) \text{ e)} \\
| & (e \ e) \\
| & ((\text{prim } e) \ e) \\
| & x \\
| & n
\end{align*}
\]

\[
\text{prim ::= + | * | …}
\]
A type system for STLC

Type rules are written in natural-deduction style
(Like our big-step operational semantics.)

Assumptions above the line  (No assumptions here.)

\[ n : \text{num} \]

Conclusions below the line

“We may conclude any number n has type \text{num}”

\[
e ::= (\lambda (x) \, e) \\
| \, (e \, e) \\
| \, ((\text{prim} \, e) \, e) \\
| \, x \\
| \, n \\
\]

\[
\text{prim} ::= + \, | \, * \, | \, \ldots
\]
Variable Lookup

We assume a **typing environment** which maps variables to their types

\[ \Gamma(x) = t \]

If \( x \) maps to type \( t \) in \( \Gamma \), we may conclude that \( x \) has type \( t \) under the type environment \( \Gamma \)
Const revisited…

“We may conclude any constant $n$ is of type `num` under any typing environment.”

\[
\Gamma \vdash n : \text{num}
\]
Functions...

If you conclude that e has type t' with Gamma plus assuming x has type t,…

\[ \Gamma[x \mapsto t] \vdash e : t' \]

\[ \Gamma \vdash (\lambda (x : t) \ e) : t \to t' \quad \text{Lam} \]

Then you can conclude that the entire lambda has type t \to t'.
Functions...

If you conclude that e has type t’ with Gamma plus assuming x has type t,…

\[ \Gamma \vdash (\lambda (x : t) \ e) : t \to t' \]

Then you can conclude that the entire lambda has type t’

Note

Variables (x) must be tagged with a type (e.g., by programmer)
\[
\Gamma[x \mapsto t] \vdash e : t' \\
\Gamma \vdash (\lambda (x : t) \; e) : t \to t'
\]

\[
\begin{array}{c}
\hline
\text{(lambda (x : num) 1)} \\
\end{array}
\]
\[
\Gamma[x \mapsto t] \vdash e : t' \\
\Gamma \vdash (\lambda (x : t) \ e) : t \rightarrow t'
\]

Lam

Start with the empty environment (since this term is closed)

\[
\Gamma = \{\} \vdash (\text{lambda} \ (x : \text{num}) \ 1) : ? \rightarrow ?
\]
We suppose there are two types, $t$ and $t'$, which will make the derivation work.
We suppose there are two types, $t$ and $t'$, which will make the derivation work.

Because $x$ is tagged, it must be $\text{num}$

$$\Gamma = \{ \} \vdash (\lambda x: \text{num} \ 1) : \text{num} \to t'$$
The **Const** rule allows us to conclude $1 : \text{num}$

$$\{x \mapsto \text{num}\} \vdash 1 : t'$$

$$\Gamma = \{\} \vdash (\lambda (x : \text{num}) \, 1) : \text{num} \rightarrow t'$$

We **suppose** there are two types, $t$ and $t'$, which will make the derivation work.
\[
\Gamma = \{\} \vdash (\text{lambda } (x : \text{num}) \ 1) : \text{num} \rightarrow \text{num}
\]

So \( t' = \text{num} \)
Function Application

$\Gamma \vdash e : t \rightarrow t'$ \hspace{1cm} $\Gamma \vdash e' : t$

$\Gamma \vdash (e \ e') : t'$

App
Function Application

If (under Gamma), e has type \( t \to t' \)

And \( e' \) (its argument) has type \( t \)

\[
\Gamma \vdash e : t \to t' \quad \Gamma \vdash e' : t
\]

\[
\Gamma \vdash (e \ e') : t'
\]

Then the application of \( e \) to \( e' \) results in a \( t' \)
Our type system so far...

\[ \begin{align*}
\Gamma & \vdash n : \text{num} \\
\Gamma & \vdash x : t \\
\Gamma & \vdash e : t \rightarrow t' \\
\Gamma & \vdash e' : t \\
\Gamma & \vdash (e \ e') : t' \\
\Gamma & \vdash (\lambda (x : t) \ e) : t \rightarrow t' \\
\end{align*} \]
Almost everything! Just need builtin functions

\[
\begin{align*}
e & ::= (\lambda (x : t) e) \\
& \mid (e \ e) \\
& \mid ((\text{prim } e) \ e) \\
& \mid x \\
& \mid n
\end{align*}
\]

\[
\text{prim} ::= + \mid * \mid \ldots
\]

**Trick!** Just assume they’re part of \( \Gamma \)!

\[
\Gamma_t = \{ + : \text{num} \to \text{num} \to \text{num}, \ldots \}
\]
Practice Derivations

Write derivations of the following expressions…
\begin{align*}
\Gamma \vdash n : \text{num} & \quad \text{Const} & \Gamma \vdash x : t & \quad \text{Var} \\
\Gamma \vdash e : t \to t' & \quad \Gamma \vdash e' : t & \quad \text{App} \\
\Gamma \vdash (e e') : t' & \\
\Gamma, \{x \mapsto t\} \vdash e : t' & \quad \text{Lam} \\
\Gamma \vdash (\lambda (x : t) e) : t \to t' & \\
\end{align*}

(((\lambda (x : \text{int}) x) \ 1)
\[((\lambda (f : \text{num} \rightarrow \text{num}) (f \ 1)) \ (\lambda (x : \text{num}) \ x))\]
Typability in STLC

Not all terms can be given types…

\((\lambda (f : \text{num} \to \text{num}) \ (f \ f))\)

It is impossible to write a derivation for the above term!

f is num->num but would need to be num!
Typability

Not all terms can be given types...

$$(((\lambda (f) (f \ f)) \ (\lambda (f) (f \ f))))$$

It is **impossible** to write a derivation for $\Omega$!

Consider what would happen if $f$ were:
- num $\rightarrow$ num
- (num $\rightarrow$ num) $\rightarrow$ num

Always just out of reach...
\[(\lambda \ (f : \text{num} \rightarrow \text{num} \rightarrow \text{num}) \ (((f \ 2) \ 3) \ 4))\]
\[(((\lambda \ (f : \text{num} \rightarrow \text{num}) \ f) \ (\lambda \ (x:\text{num}) \ (\lambda \ (x:\text{num}) \ x)))\]
Type Checking

Type checking: verifying the derivation of a fully-typed term

\((\lambda (x:\text{num}) x:\text{num}) : \text{num} \rightarrow \text{num})\)

Notice that each subterm is assigned a “full” type
Type checking tells us which rules we must apply if there is to be a derivation
In the case of fully-annotated STLC, there are no parts where we have to guess a type.

We can synthesize a type by looking at the annotated parameters for lambdas.

This leads us to writing a syntax-directed (i.e., structurally-recursive) type synthesizer / checker for fully-annotated STLC.

Next lecture, we will look at type inference for un-annotated STLC.
(define (synthesize-type env e)
  (match e
    ;; Literals
    [(? integer? i) 'int]
    [(? boolean? b) 'bool]
    ;; Look up a type variable in an environment
    [(? symbol? x) (hash-ref env x)]
    ;; Lambda w/ annotation
    [`(lambda (,x : ,A) ,e)
     `(,A -> ,(synthesize-type (hash-set env x A) e))]
    ;; Arbitrary expression
    [`(,e : ,t) (let ([e-t (synthesize-type env e)])
                 (if (equal? e-t t)
                     t
                     (error (format "types ~a and ~a are different" e-t t))))]
    ;; Application
    [`(,e1 ,e2)
     (match (synthesize-type env e1)
        [`(,A -> ,B)
         (let ([t-2 (synthesize-type env e2)])
          (if (equal? t-2 A)
            B
            (error (format "types ~a and ~a are different" A t-2))))])])
Type Inference

Allows us to leave some placeholder variables that will be “filled in later”

$$((\lambda \ (x:t) \ x:t') \ : \ \text{num} \rightarrow \ \text{num})$$

The num->num type then forces $t = \text{num}$ and $t' = \text{num}$
Type Inference

Type inference can fail, too…

\((\lambda (x) (\lambda (y:\text{num->num}) ((+ (x y)) x)))\)

No possible type for x! Used as fn and arg to +
Type Inference has been of interest (research and practical) for many years

It allows you to write **untyped** programs (much less painful!) and automatically *synthesize* a type for you—as long as the type exists (catch your mistakes)

\[
(\lambda (f) (((f\ 2)\ 3)\ 4))
\]

**Type inference**

\[
(\lambda (f:\ \text{num} \rightarrow \text{num} \rightarrow \text{num}) (((f\ 2)\ 3)\ 4))
\]

Type inference can be seen as enumerating all possible type assignments to infer a valid typing. You can think of it as solving the equation:

\[
\exists T. \ (\lambda (f:\ T) (((f\ 2)\ 3)\ 4))
\]
How hard is this problem (tractability)?

Type inference can be seen as enumerating all possible type assignments to infer a valid typing. You can think of it as solving the equation:

$$\exists T. (\lambda (f : T) (((f 2) 3) 4))$$

There are an infinite number of possible T (e.g., int, bool, int->int, bool->bool, …) that we could check, in principle

So it is not obvious that this is a terminating process. But: humans almost always write “reasonable” types:

$$((a \to ((a \to b) \to ((a \to b) \to (b \to c)))) \to \ldots$$ is possible but uncommon

We will see next lecture that a procedure exists which finds a typing, if a typing exists. This relies on unification (a principle from logic programming)
Extending STLC…

e ::= (lambda (x) e)  
   | (e e)  
   | ((prim e) e)  
   | x  
   | n

prim ::= + | * | ...

Let’s add if, and, or
Extending STLC...

e ::= (lambda (x) e)  
    | (e e) 
    | ((prim e) e) 
    | (if e e e) 
    | (and e e) 
    | (or e e) 
    | x 
    | n | #t | #f

prim ::= + | * | ...

Now we need typing rules for if!
If needs guard to be a boolean...

Shouldn’t be valid for guard to be, e.g., \((+ 1 2)\)

\[(\text{if guard t f})\]
If needs guard to be a boolean...

Shouldn’t be valid for guard to be, e.g., (+ 1 2)

\[
\begin{align*}
\text{(if guard} & \quad \text{t} \\
\text{f) }
\end{align*}
\]

\[
\begin{align*}
\Gamma \vdash e_g : \text{bool} & \quad \Gamma \vdash e_t : t & \quad \Gamma \vdash e_f : t \\
\hline
\Gamma \vdash (\text{if } e_g e_t e_f) : t
\end{align*}
\]
If needs guard to be a boolean...

Shouldn’t be valid for guard to be, e.g., (+ 1 2)

\[(\text{if} \ \text{guard} \ \text{t} \ \text{f})\]

\[\Gamma \vdash e_g : \text{bool} \quad \Gamma \vdash e_t : t \quad \Gamma \vdash e_f : t\]

\[\Gamma \vdash (\text{if} \ e_g \ e_t \ e_f) : t\]
Exercise

Can you come up with the type rules for and/or?

\((\text{and } \ e_1 \ e_2)\)
\[ \Gamma \vdash e_1 : \text{bool} \quad \Gamma \vdash e_2 : \text{bool} \]

\[ \Gamma \vdash (\text{and} \ e_1 \ e_2) : \text{bool} \]
Completeness of STLC

- **Incomplete**: Reasonable functions we can’t write in STLC
  - E.g., any program using recursion
- Several useful **extensions** to STLC
  - **Fix operator** to allow typing recursive functions
  - **Algebraic data types** to type structures
  - **Recursive types** to allow typing recursive structures
  - `tree = Leaf (int) | Node(int,tree,tree)`
Typing the Y Combinator

\[
\Gamma \vdash f : t \rightarrow t \\
\Gamma \vdash (Yf) : t
\]

The “real” solution is quite nontrivial—we need recursive types, which may be formalized in a variety of ways
- We will not cover recursive types in this lecture, I am happy to offer pointers

Our hacky solution works in practice, but is not sound in general
- More precisely, the logic induced by the type system is no longer sound
Typing the Y Combinator

Think of how this would look for \texttt{fib}

\[
\frac{\Gamma \vdash f : t \rightarrow t}{\Gamma \vdash (Yf) : t}
\]

\texttt{(let ([fib (\ Y (\ (f) (\ (x) 
(if (= x 0) 
1 
(* x (fib (- x 1))))))))])}

What would \(t\) be here?
Error States

A program steps to an **error state** if its evaluation reaches a point where the program has not produced a value, and yet cannot make progress.

\[
\text{((+ 1) (λ (x) x))}
\]

Gets “stuck” because + can’t operate on λ
Error States

A program steps to an error state if its evaluation reaches a point where the program has not produced a value, and yet cannot make progress

$$\text{(++) } (\lambda (x) \ x))$$

Gets “stuck” because + can’t operate on \(\lambda\)

(Note that this term is not typable for us!)
Soundness

A type system is **sound** if no typable program will ever evaluate to an error state

“Well typed programs cannot go wrong.” — Milner

(You can **trust** the type checker!)
Proving Type Soundness

**Theorem:** if $e$ has some type derivation, then it will evaluate to a value.

Relies on two lemmas

- **Progress**
  
  If $e$ typable, then it is either a value or can be further reduced

- **Preservation**
  
  If $e$ has type $t$, any reduction will result in a term of type $t$
“Proofs as Programs”

A significant amount of interest has been given to programming languages which use **powerful type systems** to write programs alongside a proof of the program’s correctness.

Imagine how nice it would be to write a **completely-formally-verified** program—no bugs ever again!
How does this work?

These systems interpret programs as theorems in higher-order logics (calculus of constructions, etc...)

Unfortunately, no free lunch: this makes the type system way more complicated in practical settings

We will see a taste of the inspiration for these systems, by reflecting on STLC’s expressivity
Intuitionistic Propositional Logic

Constructive logic variant of traditional propositional (boolean) logic

Proofs in (intuitionistic) propositional logic are built from natural-deduction rules, including introduction and elimination rules

- **Assumption**: \( \Gamma, P \vdash P \)
- **Conjunction Introduction**: \( \Gamma \vdash \phi \) \( \Gamma \vdash \psi \) \( \Gamma \vdash \phi \land \psi \)
- **Conjunction Left-Elimination**: \( \Gamma \vdash \phi \land \psi \) \( \Gamma \vdash \phi \)
- **Conjunction Right-Elimination**: \( \Gamma \vdash \phi \land \psi \) \( \Gamma \vdash \phi \)

Implication in IPL

Implication is performed by *introducing-then-discharging*

“If you can prove $\psi$ by assuming $\phi$, then you can prove $\phi \Rightarrow \psi$”

Sometimes called the *deduction theorem*

“If you have a proof of $\phi \Rightarrow \psi$, and a proof of $\phi$, then you can have a proof of $\psi$”

Sometimes called *modus ponens*
Proving \( P \Rightarrow (Q \Rightarrow P) \)

Assumption \[ Q, P \vdash P \]

\Rightarrow \text{Intros} \[
Q \vdash Q \Rightarrow P
\]

\Rightarrow \text{Intros} \[
P \vdash (P \Rightarrow (Q \Rightarrow P))
\]

Start with a goal and then grow a proof according to the rules
Small Point: Proving $P \Rightarrow (Q \Rightarrow Q)$

Should be a simple fix

Unfortunately, our assumption rule forbids this:

To fix this, we typically add structural rules to allow identifying contexts under reorderings. Some “substructural” logics (linear, affine) explicitly restrict this for particular uses (tracking resources, etc…).
Curry-Howard-Isomorphism

Every well-typed STLC term is a proof of a theorem in intuitionistic propositional logic

\( \lambda (x : \text{int}) \ x \) : \text{int} -> \text{int}  

Can be interpreted as “P implies P” (P \Rightarrow P, more properly \text{int} \Rightarrow \text{int})

\( \lambda (x : \text{int}) \ (\lambda (y : \text{bool}) \ x)) \) : \text{int} -> (\text{bool} -> \text{int})

Can be interpreted “P \Rightarrow (Q \Rightarrow P)”
The key idea is to realize that the typing derivation for STLC precisely mirrors the deductive rules of IPL.

\[
\frac{x \mapsto t \in \Gamma}{\Gamma \vdash x : t} \quad \text{Var}
\]

\[
\frac{\Gamma \vdash e : t \rightarrow t' \quad \Gamma \vdash e' : t}{\Gamma \vdash (e 
  e') : t'} \quad \text{App}
\]

\[
\frac{\Gamma \vdash (\lambda (x : t) \ e) : t \rightarrow t'}{\Gamma \vdash \{x \mapsto t\} \vdash e : t'} \quad \text{Lam}
\]

\[
\frac{\Gamma \vdash \phi \Rightarrow \psi \quad \Gamma \vdash \phi}{\Gamma \vdash \psi} \quad \Rightarrow I
\]

\[
\frac{\Gamma \vdash \phi \Rightarrow \psi}{\Gamma \vdash \phi \Rightarrow \psi} \quad \Rightarrow E
\]

\[
\frac{\Gamma, \phi \vdash \psi}{\Gamma \vdash \phi \Rightarrow \psi} \quad \text{Assumption}
\]
This means that every proof tree for STLC can be **trivially-mapped** to a proof tree in IPL. I.e., if \((e : t)\) is typeable in STLC, the theorem \(t\) holds in IPL by construction of the proof built using this mapping.

\[
\begin{align*}
\frac{x \mapsto t \in \Gamma}{\Gamma \vdash x : t} & \quad \text{Var} \\
\frac{\Gamma \vdash e : t \rightarrow t' \quad \Gamma \vdash e' : t}{\Gamma \vdash (e \: e') : t'} & \quad \text{App} \\
{\Gamma, \{x \mapsto t\} \vdash e : t'} & \quad \text{Lam} \\
\frac{\Gamma \vdash (\lambda (x : t) \: e) : t \rightarrow t'}{\Gamma \vdash (e \: e') : t'} & \quad \text{⇒I} \\
\frac{\Gamma \vdash \phi \Rightarrow \psi \quad \Gamma \vdash \phi}{\Gamma \vdash \psi} & \quad \text{⇒E} \\
\frac{\Gamma, P \vdash P}{\Gamma, P \vdash P} & \quad \text{Assumption}
\end{align*}
\]

\[\text{Γ, } P \vdash P\]
A family of logics / type systems

The Curry-Howard Isomorphism is a principle we can use to interpret either type systems or constructive logics
- (Always constructive logics because structural type systems are fully-materialized, structured proofs)

IPL is a boring logic—it can’t say much. Expressive power is limited to propositional logic

To prove interesting theorems, we want to say things like:
\[ \forall (l : \text{list } A) : \{l' : \text{sorted } l' \land \forall x. (\text{member } l \ x) \Rightarrow (\text{member } l' \ x)\} \]
- For all input lists l
- The output is a list l’, along with a proof that:
  - (a) l’ is sorted (specified elsewhere)
  - (b) every member of l is also a member of l’
- Any issues?
  - (Maybe we also want to assert length is the same?)
Dependent Type Systems

We can construct type systems / programming languages where terms can be of type (something like)

\[ \forall (l: \text{list } A) : \{l' : \text{sorted } l' \land \forall (x : A). (\text{member } l \ x) \Rightarrow (\text{member } l' \ x)\} \]

These are called dependent types, because types can depend on values
- This allows expressing that \( l' \) is sorted
- Unfortunately, these type systems are way more complicated
- Worse, even type checking may be undecidable (in general)

Precise formalization of these systems is beyond the scope of this class
A huge family of languages have popped up to implement dependent type systems and subsequently enable “fully-verified” programming.

They hit a variety of expressivity points. The fundamental trade off is: (a) expressivity vs. (b) automation.

Highly-expressive systems require you to write all the proofs yourself, and a lot of manual annotation (potentially).

```
prove(M,I) := append(G,[|R|],M), \member(_,I),
append(G,R,S), prove([I],[\{\text{-1}\}S],(I)),
prove(I,\ldots).
prove([|L|G],N,P,I) := \text{-}(N+1), \text{\text{-}L}\rightarrow \text{\text{-}(member(G,P));}
append(G,[|R|],N), copy_term(I,Z), append(A,[|R|],E),
append(A,B,F), (B=E \rightarrow append(B,Z,S)), length(P,H), k<1,
append(A,[|D|Q],S)), prove(F,B,\{\text{-}1\}P),I), prove(C,N,P,I).``
Here I give an Agda definition for products
Explicit Theorem Proving / Hole-Based Synth

Agda will tell me what I need to fill in, allows me to use “holes” and then helps me hunt for a working proof.
Some systems provide logic-programming (i.e., proof search) to help assist users
- CHI tells us that proof search is tantamount to program synthesis
- Here I use Coq’s “intuition” tactic to automatically construct a proof for me

(Using Coq to prove $P \Rightarrow Q \Rightarrow P$; left: using the “intuition” tactic, right: printing the proof term)
Other systems for dependent type synthesis

The more expressive the type theory, the more work is required to build proofs.

Some systems translate proof obligations into formulas which are then sent to SMT solvers (solves goals in first-order logic, such as Z3)

This can partially automate many otherwise-tricky proofs—in certain situations

F* based on this idea, but other proof search approaches exist (Idris, etc...)