Haskell-style type classes with Isabelle/Isar

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Abstract

This tutorial introduces Isar type classes, which are a convenient mechanism for organizing specifications. Essentially, they combine an operational aspect (in the manner of Haskell) with a logical aspect, both managed uniformly.

1 Introduction

Type classes were introduced by Wadler and Blott [9] into the Haskell language to allow for a reasonable implementation of overloading[1]. As a canonical example, a polymorphic equality function $eq :: \alpha \Rightarrow \alpha \Rightarrow bool$ which is overloaded on different types for α, which is achieved by splitting introduction of the *eq* function from its overloaded definitions by means of *class* and *instance* declarations: [2]

$$
\begin{array}{l}
\textit{class eq where} \\
\quad eq :: \alpha \Rightarrow \alpha \Rightarrow bool \\[4pt]
\textit{instance nat} :: \textit{eq where} \\
\quad eq\ 0\ 0 = \textit{True} \\
\quad eq\ 0\ \text{-} = \textit{False} \\
\quad eq\ \text{-}\ 0 = \textit{False} \\
\quad eq\ (Suc\ n)\ (Suc\ m) = eq\ n\ m \\[4pt]
\textit{instance}\ (\alpha{::}eq,\ \beta{::}eq)\ pair :: \textit{eq where} \\
\quad eq\ (x1,\ y1)\ (x2,\ y2) = eq\ x1\ x2 \wedge eq\ y1\ y2 \\[4pt]
\textit{class ord extends eq where} \\
\quad less\text{-}eq :: \alpha \Rightarrow \alpha \Rightarrow bool \\
\quad less :: \alpha \Rightarrow \alpha \Rightarrow bool
\end{array}
$$

Type variables are annotated with (finitely many) classes; these annotations are assertions that a particular polymorphic type provides definitions for overloaded functions.

Indeed, type classes not only allow for simple overloading but form a generic calculus, an instance of order-sorted algebra [6, 7, 10].

From a software engineering point of view, type classes roughly correspond to interfaces in object-oriented languages like Java; so, it is naturally desirable that type classes do not only provide functions (class parameters) but also state specifications implementations must obey. For example, the *class eq* above could be given the following specification, demanding that *class eq* is an equivalence relation obeying reflexivity, symmetry and transitivity:

$$
\begin{array}{l}
\textit{class eq where} \\
\quad eq :: \alpha \Rightarrow \alpha \Rightarrow bool \\
\textit{satisfying}
\end{array}
$$

[1] throughout this tutorial, we are referring to classical Haskell 1.0 type classes, not considering later additions in expressiveness

[2] syntax here is a kind of isabellized Haskell

refl: $eq\ x\ x$
sym: $eq\ x\ y \longleftrightarrow eq\ x\ y$
trans: $eq\ x\ y \wedge eq\ y\ z \longrightarrow eq\ x\ z$

From a theoretical point of view, type classes are lightweight modules; Haskell type classes may be emulated by SML functors [1]. Isabelle/Isar offers a discipline of type classes which brings all those aspects together:

1. specifying abstract parameters together with corresponding specifications,
2. instantiating those abstract parameters by a particular type
3. in connection with a "less ad-hoc" approach to overloading,
4. with a direct link to the Isabelle module system: locales [4].

Isar type classes also directly support code generation in a Haskell like fashion. Internally, they are mapped to more primitive Isabelle concepts [3].

This tutorial demonstrates common elements of structured specifications and abstract reasoning with type classes by the algebraic hierarchy of semigroups, monoids and groups. Our background theory is that of Isabelle/HOL [8], for which some familiarity is assumed.

2 A simple algebra example

2.1 Class definition

Depending on an arbitrary type α, class *semigroup* introduces a binary operator $(\otimes)$ that is assumed to be associative:

class *semigroup* =
  fixes *mult* $:: \alpha \Rightarrow \alpha \Rightarrow \alpha$ (**infixl** $\otimes$ 70)
  assumes *assoc*: $(x \otimes y) \otimes z = x \otimes (y \otimes z)$

This **class** specification consists of two parts: the *operational* part names the class parameter (**fixes**), the *logical* part specifies properties on them (**assumes**). The local **fixes** and **assumes** are lifted to the theory toplevel, yielding the global parameter *mult* $:: \alpha{::}semigroup \Rightarrow \alpha \Rightarrow \alpha$ and the global theorem *semigroup.assoc*: $\bigwedge x\ y\ z :: \alpha{::}semigroup.\ (x \otimes y) \otimes z = x \otimes (y \otimes z)$.

2.2 Class instantiation

The concrete type *int* is made a *semigroup* instance by providing a suitable definition for the class parameter $(\otimes)$ and a proof for the specification of *assoc*. This is accomplished by the **instantiation** target:

instantiation *int* :: *semigroup*
begin

definition
  mult-int-def: $i \otimes j = i + (j{::}int)$

instance proof
  fix $i\ j\ k$ :: *int* **have** $(i + j) + k = i + (j + k)$ **by** *simp*
  then show $(i \otimes j) \otimes k = i \otimes (j \otimes k)$
    unfolding *mult-int-def* **.**
qed

end

instantiation defines class parameters at a particular instance using common specification tools (here, **definition**). The concluding **instance** opens a proof that the given parameters actually conform to the class specification. Note that the first proof step is the *default* method, which for such instance proofs maps to the *intro-classes* method. This reduces an instance judgement to the relevant primitive proof goals; typically it is the first method applied in an instantiation proof.

From now on, the type-checker will consider *int* as a *semigroup* automatically, i.e. any general results are immediately available on concrete instances.

Another instance of *semigroup* yields the natural numbers:

instantiation *nat* :: *semigroup*
begin

primrec *mult-nat* **where**
  $(0{::}nat) \otimes n = n$
  $\mid Suc\ m \otimes n = Suc\ (m \otimes n)$

instance proof
  fix $m\ n\ q$ :: *nat*
  show $m \otimes n \otimes q = m \otimes (n \otimes q)$
    by (*induct m*) *auto*

qed

end

Note the occurence of the name *mult-nat* in the primrec declaration; by default, the local name of a class operation f to be instantiated on type constructor κ is mangled as f-κ. In case of uncertainty, these names may be inspected using the **print-context** command or the corresponding ProofGeneral button.

2.3 Lifting and parametric types

Overloaded definitions given at a class instantiation may include recursion over the syntactic structure of types. As a canonical example, we model product semigroups using our simple algebra:

instantiation *prod* :: (*semigroup*, *semigroup*) *semigroup*
begin

definition
  mult-prod-def: $p_1 \otimes p_2 = (\mathit{fst}\ p_1 \otimes \mathit{fst}\ p_2,\ \mathit{snd}\ p_1 \otimes \mathit{snd}\ p_2)$

instance proof
  fix $p_1\ p_2\ p_3 :: \alpha{::}\mathit{semigroup} \times \beta{::}\mathit{semigroup}$
  show $p_1 \otimes p_2 \otimes p_3 = p_1 \otimes (p_2 \otimes p_3)$
    unfolding *mult-prod-def* **by** (*simp add*: *assoc*)
qed

end

Associativity of product semigroups is established using the definition of $(\otimes)$ on products and the hypothetical associativity of the type components; these hypotheses are legitimate due to the *semigroup* constraints imposed on the type components by the **instance** proposition. Indeed, this pattern often occurs with parametric types and type classes.

2.4 Subclassing

We define a subclass *monoidl* (a semigroup with a left-hand neutral) by extending *semigroup* with one additional parameter *neutral* together with its characteristic property:

class *monoidl* = *semigroup* +
  fixes *neutral* :: α (**1**)
  assumes *neutl*: **1** $\otimes$ $x = x$

Again, we prove some instances, by providing suitable parameter definitions and proofs for the additional specifications. Observe that instantiations for types with the same arity may be simultaneous:

instantiation *nat* **and** *int* :: *monoidl*
begin

definition
  neutral-nat-def: **1** = (0::*nat*)

definition
  neutral-int-def: **1** = (0::*int*)

instance proof
  fix n :: *nat*
  show **1** $\otimes$ $n = n$
    unfolding *neutral-nat-def* **by** *simp*
next
  fix k :: *int*
  show **1** $\otimes$ $k = k$
    unfolding *neutral-int-def mult-int-def* **by** *simp*
qed

end

instantiation *prod* :: (*monoidl*, *monoidl*) *monoidl*
begin

definition
  neutral-prod-def: **1** = (**1**, **1**)

instance proof
  fix p :: α::*monoidl* $\times$ β::*monoidl*
  show **1** $\otimes$ $p = p$
    unfolding *neutral-prod-def mult-prod-def* **by** (*simp add*: *neutl*)
qed

end

Fully-fledged monoids are modelled by another subclass, which does not add new parameters but tightens the specification:

class *monoid* = *monoidl* +
  assumes *neutr*: $x \otimes \mathbf{1} = x$

instantiation *nat* **and** *int* :: *monoid*
begin

instance proof
  fix n :: *nat*
  show $n \otimes \mathbf{1} = n$
    unfolding *neutral-nat-def* **by** (*induct n*) *simp-all*
next
  fix k :: *int*
  show $k \otimes \mathbf{1} = k$
    unfolding *neutral-int-def mult-int-def* **by** *simp*
qed

end

instantiation *prod* :: (*monoid*, *monoid*) *monoid*
begin

instance proof
  fix p :: α::*monoid* × β::*monoid*
  show $p \otimes \mathbf{1} = p$
    unfolding *neutral-prod-def mult-prod-def* **by** (*simp add*: *neutr*)
qed

end

To finish our small algebra example, we add a *group* class with a corresponding instance:

class *group* = *monoidl* +
  fixes *inverse* :: $\alpha \Rightarrow \alpha$ ((-÷) [1000] 999)
  assumes *invl*: $x\div \otimes x = \mathbf{1}$

instantiation *int* :: *group*
begin

definition

inverse-int-def: $i\div = - (i{::}int)$

instance proof
fix $i :: int$
have $-i + i = 0$ **by** *simp*
then show $i\div \otimes i = \mathbf{1}$
unfolding *mult-int-def neutral-int-def inverse-int-def* **.**
qed

end

3 Type classes as locales

3.1 A look behind the scenes

The example above gives an impression how Isar type classes work in practice. As stated in the introduction, classes also provide a link to Isar's locale system. Indeed, the logical core of a class is nothing other than a locale:

class *idem* $=$
fixes $f :: \alpha \Rightarrow \alpha$
assumes *idem*: $f\ (f\ x) = f\ x$

essentially introduces the locale

locale *idem* $=$
fixes $f :: \alpha \Rightarrow \alpha$
assumes *idem*: $f\ (f\ x) = f\ x$

together with corresponding constant(s):

consts $f :: \alpha \Rightarrow \alpha$

The connection to the type system is done by means of a primitive type class

classes *idem* $<$ *type*

together with a corresponding interpretation:

interpretation *idem-class*:
idem $f :: (\alpha{::}idem) \Rightarrow \alpha$

This gives you the full power of the Isabelle module system; conclusions in locale *idem* are implicitly propagated to class *idem*.

3.2 Abstract reasoning

Isabelle locales enable reasoning at a general level, while results are implicitly transferred to all instances. For example, we can now establish the *left-cancel* lemma for groups, which states that the function $(x \otimes)$ is injective:

lemma (**in** *group*) *left-cancel*: $x \otimes y = x \otimes z \longleftrightarrow y = z$
proof
  assume $x \otimes y = x \otimes z$
  then have $x\div \otimes (x \otimes y) = x\div \otimes (x \otimes z)$ **by** *simp*
  then have $(x\div \otimes x) \otimes y = (x\div \otimes x) \otimes z$ **using** *assoc* **by** *simp*
  then show $y = z$ **using** *neutl* **and** *invl* **by** *simp*
next
  assume $y = z$
  then show $x \otimes y = x \otimes z$ **by** *simp*
qed

Here the "**in** *group*" target specification indicates that the result is recorded within that context for later use. This local theorem is also lifted to the global one *group.left-cancel*: $\bigwedge x\ y\ z :: \alpha{::}group.\ x \otimes y = x \otimes z \longleftrightarrow y = z$. Since type *int* has been made an instance of *group* before, we may refer to that fact as well: $\bigwedge x\ y\ z :: int.\ x \otimes y = x \otimes z \longleftrightarrow y = z$.

3.3 Derived definitions

Isabelle locales are targets which support local definitions:

primrec (**in** *monoid*) *pow-nat* :: $nat \Rightarrow \alpha \Rightarrow \alpha$ **where**
  pow-nat $0\ x = \mathbf{1}$
  | *pow-nat* $(Suc\ n)\ x = x \otimes$ *pow-nat* $n\ x$

If the locale *group* is also a class, this local definition is propagated onto a global definition of *pow-nat* :: $nat \Rightarrow \alpha{::}monoid \Rightarrow \alpha{::}monoid$ with corresponding theorems
  pow-nat $0\ x = \mathbf{1}$
  pow-nat $(Suc\ n)\ x = x \otimes$ *pow-nat* $n\ x$.
As you can see from this example, for local definitions you may use any specification tool which works together with locales, such as Krauss's recursive function package [5].

3.4 A functor analogy

We introduced Isar classes by analogy to type classes in functional programming; if we reconsider this in the context of what has been said about type classes and locales, we can drive this analogy further by stating that type classes essentially correspond to functors that have a canonical interpretation as type classes. There is also the possibility of other interpretations. For example, *list*s also form a monoid with *append* and [] as operations, but it seems inappropriate to apply to lists the same operations as for genuinely algebraic types. In such a case, we can simply make a particular interpretation of monoids for lists:

interpretation *list-monoid*: *monoid append* []
  proof qed *auto*

This enables us to apply facts on monoids to lists, e.g. [] @ $x = x$.

When using this interpretation pattern, it may also be appropriate to map derived definitions accordingly:

primrec *replicate* :: *nat* ⇒ α *list* ⇒ α *list* **where**
  replicate 0 - = []
  | *replicate* (*Suc n*) *xs* = *xs* @ *replicate n xs*

interpretation *list-monoid*: *monoid append* [] **where**
  monoid.pow-nat append [] = *replicate*
proof −
  interpret *monoid append* [] ..
  show *monoid.pow-nat append* [] = *replicate*
  proof
    fix *n*
    show *monoid.pow-nat append* [] *n* = *replicate n*
      by (*induct n*) *auto*
  qed
qed *intro-locales*

This pattern is also helpful to reuse abstract specifications on the *same* type. For example, think of a class *preorder*; for type *nat*, there are at least two possible instances: the natural order or the order induced by the divides relation. But only one of these instances can be used for **instantiation**; using the locale behind the class *preorder*, it is still possible to utilise the same abstract specification again using **interpretation**.

3.5 Additional subclass relations

Any *group* is also a *monoid*; this can be made explicit by claiming an additional subclass relation, together with a proof of the logical difference:

subclass (**in** *group*) *monoid*
proof
  fix x
  from *invl* **have** $x\div \otimes x = \mathbf{1}$ **by** *simp*
  with *assoc* [*symmetric*] *neutl invl* **have** $x\div \otimes (x \otimes \mathbf{1}) = x\div \otimes x$ **by** *simp*
  with *left-cancel* **show** $x \otimes \mathbf{1} = x$ **by** *simp*
qed

The logical proof is carried out on the locale level. Afterwards it is propagated to the type system, making *group* an instance of *monoid* by adding an additional edge to the graph of subclass relations (figure 1).

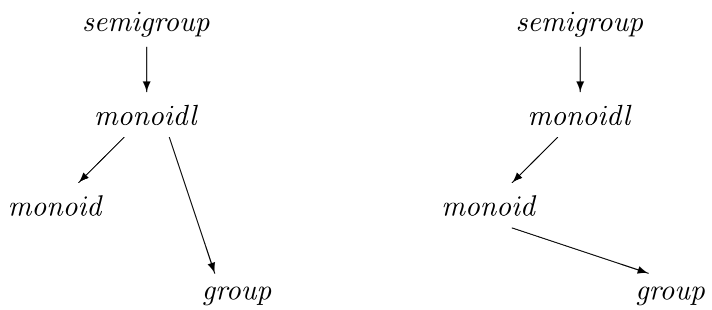

Figure 1: Subclass relationship of monoids and groups: before and after establishing the relationship *group* $\subseteq$ *monoid*; transitive edges are left out.

For illustration, a derived definition in *group* using *pow-nat*

definition (**in** *group*) *pow-int* :: *int* $\Rightarrow \alpha \Rightarrow \alpha$ **where**
  pow-int k x = (*if k* >= 0
    then pow-nat (*nat k*) *x*
    else (*pow-nat* (*nat* (− *k*)) *x*)÷)

yields the global definition of *pow-int* :: *int* $\Rightarrow \alpha$::*group* $\Rightarrow \alpha$::*group* with the corresponding theorem *pow-int k x* = (*if* $0 \leq k$ *then pow-nat* (*nat k*) *x else* (*pow-nat* (*nat* (− *k*)) *x*)÷).

3.6 A note on syntax

As a convenience, class context syntax allows references to local class operations and their global counterparts uniformly; type inference resolves ambiguities. For example:

context *semigroup*
begin

term $x \otimes y$ — example 1
term $(x::nat) \otimes y$ — example 2

end

term $x \otimes y$ — example 3

Here in example 1, the term refers to the local class operation *mult* $[\alpha]$, whereas in example 2 the type constraint enforces the global class operation *mult* $[nat]$. In the global context in example 3, the reference is to the polymorphic global class operation *mult* $[?\alpha :: semigroup]$.

4 Further issues

4.1 Type classes and code generation

Turning back to the first motivation for type classes, namely overloading, it is obvious that overloading stemming from **class** statements and **instantiation** targets naturally maps to Haskell type classes. The code generator framework [2] takes this into account. If the target language (e.g. SML) lacks type classes, then they are implemented by an explicit dictionary construction. As example, let's go back to the power function:

definition *example* :: *int* **where**
  example = *pow-int* 10 (-2)

This maps to Haskell as follows:

```
module Example where {

data Nat = Zero_nat | Suc Nat;

nat_aux :: Integer -> Nat -> Nat;
nat_aux i n = (if i <= 0 then n else nat_aux (i - 1) (Suc n));
```

```
nat :: Integer -> Nat;
nat i = nat_aux i Zero_nat;

class Semigroup a where {
  mult :: a -> a -> a;
};

class (Semigroup a) => Monoidl a where {
  neutral :: a;
};

class (Monoidl a) => Monoid a where {
};

class (Monoid a) => Group a where {
  inverse :: a -> a;
};

pow_nat :: forall a. (Monoid a) => Nat -> a -> a;
pow_nat Zero_nat x = neutral;
pow_nat (Suc n) x = mult x (pow_nat n x);

pow_int :: forall a. (Group a) => Integer -> a -> a;
pow_int k x =
  (if 0 <= k then pow_nat (nat k) x
    else inverse (pow_nat (nat (negate k)) x));

mult_int :: Integer -> Integer -> Integer;
mult_int i j = i + j;

neutral_int :: Integer;
neutral_int = 0;

instance Semigroup Integer where {
  mult = mult_int;
};

instance Monoidl Integer where {
  neutral = neutral_int;
};

instance Monoid Integer where {
};

inverse_int :: Integer -> Integer;
inverse_int i = negate i;

instance Group Integer where {
  inverse = inverse_int;
};

example :: Integer;
example = pow_int 10 (-2);

}
```

The code in SML has explicit dictionary passing:

```
structure Example : sig
  datatype nat = Zero_nat | Suc of nat
  val nat_aux : IntInf.int -> nat -> nat
  val nat : IntInf.int -> nat
```

```
  type 'a semigroup
  val mult : 'a semigroup -> 'a -> 'a -> 'a
  type 'a monoidl
  val semigroup_monoidl : 'a monoidl -> 'a semigroup
  val neutral : 'a monoidl -> 'a
  type 'a monoid
  val monoidl_monoid : 'a monoid -> 'a monoidl
  type 'a group
  val monoid_group : 'a group -> 'a monoid
  val inverse : 'a group -> 'a -> 'a
  val pow_nat : 'a monoid -> nat -> 'a -> 'a
  val pow_int : 'a group -> IntInf.int -> 'a -> 'a
  val mult_int : IntInf.int -> IntInf.int -> IntInf.int
  val neutral_int : IntInf.int
  val semigroup_int : IntInf.int semigroup
  val monoidl_int : IntInf.int monoidl
  val monoid_int : IntInf.int monoid
  val inverse_int : IntInf.int -> IntInf.int
  val group_int : IntInf.int group
  val example : IntInf.int
end = struct

datatype nat = Zero_nat | Suc of nat;

fun nat_aux i n =
  (if IntInf.<= (i, (0 : IntInf.int)) then n
    else nat_aux (IntInf.- (i, (1 : IntInf.int))) (Suc n));

fun nat i = nat_aux i Zero_nat;

type 'a semigroup = {mult : 'a -> 'a -> 'a};
val mult = #mult : 'a semigroup -> 'a -> 'a -> 'a;

type 'a monoidl = {semigroup_monoidl : 'a semigroup, neutral : 'a};
val semigroup_monoidl = #semigroup_monoidl : 'a monoidl -> 'a semigroup;
val neutral = #neutral : 'a monoidl -> 'a;

type 'a monoid = {monoidl_monoid : 'a monoidl};
val monoidl_monoid = #monoidl_monoid : 'a monoid -> 'a monoidl;

type 'a group = {monoid_group : 'a monoid, inverse : 'a -> 'a};
val monoid_group = #monoid_group : 'a group -> 'a monoid;
val inverse = #inverse : 'a group -> 'a -> 'a;

fun pow_nat A_ Zero_nat x = neutral (monoidl_monoid A_)
  | pow_nat A_ (Suc n) x =
    mult ((semigroup_monoidl o monoidl_monoid) A_) x (pow_nat A_ n x);

fun pow_int A_ k x =
  (if IntInf.<= ((0 : IntInf.int), k)
    then pow_nat (monoid_group A_) (nat k) x
    else inverse A_ (pow_nat (monoid_group A_) (nat (IntInf.~ k)) x));

fun mult_int i j = IntInf.+ (i, j);

val neutral_int : IntInf.int = (0 : IntInf.int);

val semigroup_int = {mult = mult_int} : IntInf.int semigroup;

val monoidl_int =
  {semigroup_monoidl = semigroup_int, neutral = neutral_int} :
  IntInf.int monoidl;

val monoid_int = {monoidl_monoid = monoidl_int} : IntInf.int monoid;
```

```
fun inverse_int i = IntInf.~ i;

val group_int = {monoid_group = monoid_int, inverse = inverse_int} :
  IntInf.int group;

val example : IntInf.int =
  pow_int group_int (10 : IntInf.int) (~2 : IntInf.int);

end; (*struct Example*)
```

In Scala, implicts are used as dictionaries:

```
object Example {

abstract sealed class nat
final case object Zero_nat extends nat
final case class Suc(a: nat) extends nat

def nat_aux(i: BigInt, n: nat): nat =
  (if (i <= BigInt(0)) n else nat_aux(i - BigInt(1), Suc(n)))

def nat(i: BigInt): nat = nat_aux(i, Zero_nat)

trait semigroup[A] {
  val `Example.mult`: (A, A) => A
}
def mult[A](a: A, b: A)(implicit A: semigroup[A]): A =
  A.`Example.mult`(a, b)

trait monoidl[A] extends semigroup[A] {
  val `Example.neutral`: A
}
def neutral[A](implicit A: monoidl[A]): A = A.`Example.neutral`

trait monoid[A] extends monoidl[A] {
}

trait group[A] extends monoid[A] {
  val `Example.inverse`: A => A
}
def inverse[A](a: A)(implicit A: group[A]): A = A.`Example.inverse`(a)

def pow_nat[A : monoid](xa0: nat, x: A): A = (xa0, x) match {
  case (Zero_nat, x) => neutral[A]
  case (Suc(n), x) => mult[A](x, pow_nat[A](n, x))
}

def pow_int[A : group](k: BigInt, x: A): A =
  (if (BigInt(0) <= k) pow_nat[A](nat(k), x)
    else inverse[A](pow_nat[A](nat((- k)), x)))

def mult_int(i: BigInt, j: BigInt): BigInt = i + j

def neutral_int: BigInt = BigInt(0)

implicit def semigroup_int: semigroup[BigInt] = new semigroup[BigInt] {
  val `Example.mult` = (a: BigInt, b: BigInt) => mult_int(a, b)
}

implicit def monoidl_int: monoidl[BigInt] = new monoidl[BigInt] {
  val `Example.neutral` = neutral_int
  val `Example.mult` = (a: BigInt, b: BigInt) => mult_int(a, b)
```

```
  }

  implicit def monoid_int: monoid[BigInt] = new monoid[BigInt] {
    val `Example.neutral` = neutral_int
    val `Example.mult` = (a: BigInt, b: BigInt) => mult_int(a, b)
  }

  def inverse_int(i: BigInt): BigInt = (- i)

  implicit def group_int: group[BigInt] = new group[BigInt] {
    val `Example.inverse` = (a: BigInt) => inverse_int(a)
    val `Example.neutral` = neutral_int
    val `Example.mult` = (a: BigInt, b: BigInt) => mult_int(a, b)
  }

  def example: BigInt = pow_int[BigInt](BigInt(10), BigInt(- 2))

  } /* object Example */
```

4.2 Inspecting the type class universe

To facilitate orientation in complex subclass structures, two diagnostics commands are provided:

print-classes print a list of all classes together with associated operations etc.

class-deps visualizes the subclass relation between all classes as a Hasse diagram.

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